Building Extensible and Secure Networks

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by

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Abstract

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In this dissertation, we present a network design called Rule-Based Forwarding (RBF) that provides flexible and policy-compliant forwarding. Our proposal centers around a new architectural concept: that of packet *rules*. A rule is a simple if-then-else construct that describes the manner in which the network should – or should not – forward packets. A packet identifies the rule by which it is to be forwarded and routers forward each packet in accordance with its associated rule. On one hand, rules are flexible, as they can explicitly specify paths and invoke packet processing inside the network. This enables RBF to support many previously proposed Internet extensions, such as explicit middleboxes, multiple paths, source routing and support for host mobility. On the other hand, rules are certified, which guarantees that packets comply with the policies of the parties forwarding them. This property also enables a more secure architecture, since unwanted packets can be dropped in the network, allowing RBF to stop denial of service (DoS) attacks. Using our prototype router implementation we show that the overhead RBF imposes is within the capabilities of modern network equipment.

We also describe how the ideas behind RBF can be used to improve access control in cloud computing, and present CloudPolice an access control mechanism implemented in hypervisors. CloudPolice scales to millions of hosts, is independent of the network topology, routing and addressing, and can specify flexible access control policies. These properties are not provided by traditional access control mechanisms, because these mechanisms were originally designed for enterprise environments that do not share the same challenges as cloud computing.

To my wife and my parents

Contents

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Chapter 1

Introduction

A central component of a network design is its forwarding architecture that determines the manner in which packets are transported between two endpoints. Today's Internet offers users a simple forwarding model: a user hands the network a packet with a destination address and the network makes a best-effort attempt to deliver the packet to the destination. Although simple, this architecture is also fairly limited and there have been repeated calls to extend the Internet's forwarding architecture for greater *flexibility*—allowing, for example, the user to select the path his packets should traverse [36, 90, 98, 104] or to specify whether packets can/should be processed by middleboxes and active routers [42, 61, 98, 101, 104].

Achieving a flexible forwarding architecture has thus been a long-held, if elusive, goal of Internet research [36, 42, 61, 86, 98, 101, 104]. Our work in this dissertation shares this goal. Our point of divergence from prior efforts starts with the observation that forwarding flexibility is inherently coupled with issues of *policy*.

Our thesis is that achieving flexibility is not just a matter of augmenting packets with more expressive forwarding directives that routers execute. Rather, in addition, *for each forwarding directive that enhances flexibility, the parties involved in forwarding should be able to set policies that constrain that directive*. By the policy of entity `A` (host, middlebox operator or ISP) we refer to the decision whether to approve or reject a forwarding directive based on `A`'s business or technical goals. By forwarding directive we refer to instructions provided by endpoints to routers and middleboxes on how to forward their packets. For example, a forwarding directive could specify that sender `S` can forward its packets through middlebox `M` before reaching destination `D`. An example of policy would be `M` refusing to accept packets from `S`.

To better illustrate our thesis, consider its application to the Internet. Since the main forwarding directive in IP is for sender `S` to send packets to destination `D`, `D` should be able to specify that the traffic from `S` should not reach it, *i.e.,* either by explicitly allowing or denying packets from `D`. Unfortunately, IP does not provide such functionality, effectively leaving the end-hosts vulnerable to DoS attacks. Unsurprisingly, this lack of functionality has been identified as one of the main security vulnerabilities of the Internet, and several

solutions have been proposed to address this limitation [37, 38, 68, 78, 108, 110].

Of course, forwarding directives and policies are only as good as the ability of the network to enforce them and to guarantee their authenticity. What complicates policy enforcement is the involvement of multiple parties in achieving the packet's flexible behavior—the network service providers along the path, potential middlebox operators and, of course, the source and destination. As such, the network must ensure that a packet's forwarding directive complies with the policies of *all* parties involved. In our previous middlebox example, the network must ensure that `M` is willing to relay packets from `S` to `D`. If `M` does not approve, the network should simply drop the packets before reaching `M`.

In this dissertation, we propose a new *rule-based* forwarding architecture, RBF, that treats flexibility and policy enforcement as equal design goals. RBF is based on a new architectural concept – that of packet *rules*. In RBF, instead of sending packets to a destination (IP) address, end-hosts send packets to a rule. Rules are created by destinations. A sender fetches the destination's rule from a DNS-like infrastructure and inserts it in the packets sent to that destination.

A rule is a simple `if-then-else` construct that describes the manner in which the network should – or should not – forward packets. For example, a destination `A` can receive packets only from source `S` using the rule:

R_A: `if(packet.source` $\neq$ `S)`
`drop packet`

Or a mobile client `B` might route certain video content through a 3rd-party transcoding proxy with:

R_B: `if(packet.URL = videowebsite.com)`
`sendPacketTo transcoderProxy`

The above examples are anecdotal (we present precise syntax Section 3 and additional examples in Section 3.3) but serve to illustrate how destinations can control and customize how the network forwards their packets in a manner not easily accommodated by current IP. In effect, with rules, a receiving host must specify both *which* packets it is willing to receive as well as *how* it wants these packets forwarded and processed by the network.

The rule-based architecture we develop offers the following properties:

1. **Rules are mandatory:** routers drop packets without rules
2. **Rules are provably authorized:** all recipients (end-hosts, middleboxes and/or routers) named in the rule must explicitly agree to receive the associated packet(s). Routers, middleboxes and end-users can verify a rule's authorization.
3. **Rules are provably safe:** rules cannot exhaust network resources (*e.g.,* cause forwarding loops) and cannot compromise routers.

4. **Rules allow flexible forwarding:** rules are a (constrained) program that allows a user to "customize" how the network forwards its packets.

The first two properties assist in policy enforcement by ensuring a packet is only forwarded if explicitly cleared by all recipients (*i.e.,* if it conforms with the policies of all recipients) specified in the rule. Since RBF defines policies on rules, any recipient will have the ultimate say on whether to accept any rule that contains forwarding directives sending packets to it. Since all forwarding directives are encoded into rules, we achieve our goal of enabling any entity affected by a forwarding directive to constrain that directive.

The third property ensures rules cannot be (mis)used to attack the network itself. As we shall show, the last property provides flexibility since users can give the network fine-grained instructions on how to handle their packets, enabling: explicit use of in-network functionality at middleboxes and routers, loose path forwarding, multipath forwarding, anycast, multicast, mobility, filtering of undesired senders/ports/protocols, recording of on-path information, *etc.* In this dissertation, we present the design, implementation and evaluation of a forwarding architecture that meets the above properties.

RBF relates to an extensive body of work on both forwarding flexibility and policy enforcement. We discuss related work in detail in Section 7 and here only note that, at a high level, we believe what distinguishes RBF is its focus on *simultaneously* supporting flexibility and the multi-party policy requirements that such flexibility implies. As we shall see, this goal leads us to a design that differs significantly from prior proposals.

We note up-front that RBF is more complex than the existing IP forwarding architecture, which is frequently cited for its simplicity. In addition, RBF relies on strong assumptions such as anti-spoofing, the existence of rule-certifying authorities and a DNS-like infrastructure to distribute rules. The gain, relative to today's IP forwarding, is significantly improved flexibility and security; we posit that the greater complexity of our solution is a perhaps inevitable consequence of this richer service model.

In this dissertation, we also explore the application of the core ideas behind RBF to cloud computing. Cloud computing is growing in importance as a new paradigm for access to computational resources. Core to cloud computing is the ability to share and multiplex resources across multiple users. Since cloud computing environments are shared by multiple users, they should ideally provide network-level access control mechanisms between users (typically called tenants). The majority of existing access control techniques used for cloud computing were originally designed for enterprise environments. However, compared to enterprises, cloud computing environments impose new challenges on access control techniques due to multi-tenancy, the growing scale and dynamicity of hosts within the cloud infrastructure, and the increasing diversity of data center network architectures. Therefore, the access control mechanisms inherited from enterprises are poorly suited for cloud environments. In fact, as cloud computing services evolve, they run into analogous security and inflexibility problems as the Internet. Clouds are expanding to very large sizes (currently, hundreds of thousands of machines) and, like the Internet, they have tens of thousands of users. The

latter also leads to growing concerns about potential DoS attacks between users inside the cloud. One of the properties offered by RBF is access control at the Internet level.

Starting from this observation, we propose CloudPolice. Similarly to RBF, in CloudPolice the access control decisions are specified by each destination on the control plane (conforming to the destination's policy) and are enforced on the data plane by having packets carry a proof of their policy compliance. On the other hand, CloudPolice takes advantage of the fact that, unlike the Internet, each cloud is owned by a single administrative domain that owns not only the network, but also the servers. In particular, cloud computing is typically a virtualized environment, where a trusted layer of software, the hypervisor, sits between each customer machine and the network. In this context, we argue that it is both sufficient and advantageous to implement access control *only* within hypervisors at end-hosts. For this reason, CloudPolice has a lower overhead and is easier to adopt when compared to RBF. In fact, cloud providers can adopt CloudPolice through a simple software update to hypervisors, without requiring any hardware upgrades or changes to applications. Our evaluation shows that CloudPolice does not impose a significant overhead on hosts.

1.1 Contributions

The contributions of this dissertation are as follows:

1. We present the design of RBF, a first network architecture that is both flexible and policy compliant. While many have proposed more flexible networks, we argue that adding more flexibility alone can, in fact, lead to less secure architectures. In addition, we argue that respecting the policy of destinations is also necessary in order to improve the security of today's networks and block DoS attacks. Furthermore, the lack of a guaranteed policy compliance property is one of the reasons that makes the adoption of new architectural proposals difficult, since, without such a property, many of the ISPs' concerns are not addressed.

2. We analyze RBF's flexibility and its security properties. We show that RBF can support a large number of desirable and previously proposed extensions to today's networks such as explicit use of in-network functionality at middleboxes and routers, loose path forwarding, multipath forwarding, anycast, multicast, mobility, filtering of undesired senders/ports/protocols, recording of on-path information, *etc.* Our security analysis shows that RBF is resistant to attacks from endpoints trying to forge, tamper or evade rules and that users cannot create malicious rules to attack the network itself.

3. We evaluate RBF's feasibility. We show that the increase in the traffic size for throughput intensive applications can be as low as 6% on average. Using a prototype software router implementation, we show that RBF's overhead compared to the state of the art software routers is minimal.

4. We present CloudPolice, a mechanism to implement access control in cloud computing inspired by RBF. Compared to the traditional access control mechanisms inherited from enterprises, CloudPolice can scale to millions of hosts, is independent of the network topology, routing and addressing, and can specify more flexible access control policies. In addition, CloudPolice is easy to adopt by cloud providers, requiring only a software update.

1.2 Organization

This dissertation is organized as follows. We start by presenting the design rationale and the overview of RBF in Section 2. We then present the details of RBF's data plane and a set of examples that can be supported by using the described data plane (Section 3). Next, we detail RBF's control plane (Section 4). Section 5 presents a security analysis of RBF and Section 6 presents details of the implementation and evaluation of RBF. We present related work in Section 7 and discuss different concerns associated with RBF and RBF's limitations in Section 8. Section 9 presents CloudPolice, a mechanism inspired by RBF to provide access control in data centers. Finally, Section 10 outlines our plans for future work and Section 11 concludes this dissertation.

Chapter 2

Design Rationale and Overview

We start with the goal of network flexibility and allowing users control over how the network processes their packets. The abstraction that perhaps best supports flexibility is simply that of a *program*, leading to an architecture where users write packet-processing programs that routers execute. This vision of code-carrying packets is, of course, the cornerstone of active networking [101, 105] and we borrow this as our starting point in designing RBF. However, as we shall see, RBF severely dials back on the full-fledged generality of the original active networks' vision to arrive at a significantly simpler and safer architecture.

Rules are thus a form of program. The challenge then is to appropriately constrain these programs/rules to ensure that they cannot harm the network or other hosts. The key insight behind RBF is that these constraints must extend along *two* dimensions. First, rules must be *safe*, *i.e.,* guaranteed not to corrupt or exhaust network resources. In addition, however, we must constrain rules to respect the *policies* of all stakeholders involved—source, destination, middleboxes and ISPs. This latter requirement is unique and yet critical to networking contexts but was under appreciated in early active networking proposals.

To address policy safety, RBF incorporates two key design decisions:

- **(D1) Layering:** we believe network operators will be unwilling to relinquish control of route discovery and computation and hence we layer RBF above current IP forwarding and do not allow rules to modify the IP-layer forwarding information base (FIB).
- **(D2) Verifiable stakeholder agreement:** we require that a rule be authorized by all entities it explicitly names (*e.g.,* destination, middleboxes or routers). This ensures agreement of the stakeholders' policies with the rule's intent; in particular it also ensures that rules cannot violate ISPs' routing policies, since providers must explicitly agree to have their routers named in rules. To achieve this property, in RBF rules are certified by trusted third parties, which in turn gather proofs of policy compliance from each of the rule participants.

To address rule safety, we impose strict constraints on rule syntax, such that safety can be verified through simple static analysis:

- **(C1) Rules cannot directly modify router state.** This avoids corruption of router state. However, this can be a limiting restriction, particularly to network operators who wish to expose in-network services such as caching or monitoring to end users. To accommodate this, RBF allows operators to deploy specialized packet-processing functions at their routers and allows rules to *invoke* these functions. Such "router-defined functions" do allow rules to update router state, but only indirectly via code installed, and hence presumably trusted, by operators. This model for router-defined functions thus represents a middle ground in the tradeoff between flexibility and safety.

- **(C2) The rule "instruction set" is limited** to only four possible *action* statements: (a) `forward` the packet to the underlying IP layer, (b) `invoke` a router-defined function, (c) `modify` the packet header and (d) `drop` the packet, plus conditionals that determine whether an action should be taken based on reading packet headers and router state. Note that there is no action that allows backward jumps across rule statements. This prevents looping or resource exhaustion at routers and ensures execution time is linear in program size.

The above constraints represent a stark departure from the rich generality of the active networks vision. Indeed, rules are more a sequence of packet steering directives, rather than a full-fledged program. The benefit is *verifiable rule and policy safety.* Moreover we find that, despite these constraints, rules suffice to express a wide variety of forwarding behaviors as we will later illustrate.

2.1 Architecture Overview

In this section we provide an overview of the main components and assumptions of the RBF architecture and the rationale behind their design. The subsequent sections of this dissertation will present the detailed design of these components.

With RBF, end-users control packet forwarding using rules. Each packet has an associated rule and contains a set of *attribute-value* pairs in its header. Upon receiving a packet, the router forwards the packet conforming to the packet's rule.

2.1.1 Distributing rules to routers

To process a packet, a router needs to know the rule associated with the packet. There are two basic approaches by which routers can obtain rules: (1) rules are carried in packets, (2) routers use an out-of-band mechanism to obtain the rules.

Carrying rules in packets frees routers from maintaining per-rule state, and implementing costly rule distribution protocols. On the other hand, this approach incurs a high overhead on the data plane as rules increase packet headers, and routers need to verify the rules to ensure their validity.

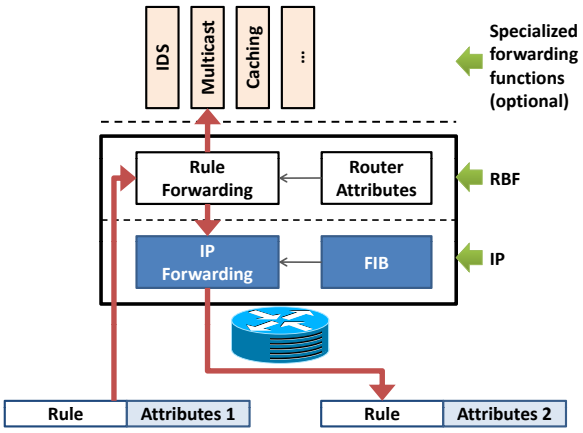


Figure 2.1: RBF router and rule forwarding

In contrast, installing rules at routers has a lower overhead on the data plane, as packets need only to carry rule identifiers instead of the rules themselves. However, the process of obtaining rules can be complex; the router can either get rules in advance, in which case it may need to store a huge number of rules or the router can download rules on demand, in which case the router may need to buffer the packets it receives until it obtains the rules for those packets. Moreover, since we do not want to have any type of "special" traffic that travels without using rules, the packets installing rules have to themselves travel on rules. This makes the rule installation process very difficult.

For these reasons, we chose to have packets carry rules. In Section 6 we evaluate the overhead of rules.

2.1.2 Applying Rules

Figure 2.1 illustrates the forwarding architecture of an RBF-enabled router. On receiving a packet, the router hands it to the *rule forwarding engine*, which processes the packet's rule. Such processing may involve reading router state that the network operator has opted to expose; we term such state *router attributes*. Examples of routing attributes may include the router's IP address, AS number, congestion level, flags indicating whether the router implements a specific functionality such as intrusion detection and so forth. Based on information in the packet header (*packet attributes*) and router attributes, the rule forwarding engine may update the value of the packet attributes (including its destination), invoke router functions, drop the packet and/or hand the packet to the underlying IP forwarding engine. Recall that for safety reasons the rule is not allowed to update router state.

2.1.3 Distributing rules to end-hosts

As discussed above, in our instantiation of RBF, rules are carried in the packets. Since the first router on the packet's path expects each packet to carry its associated rule, rules need to be inserted by senders. Furthermore, since in general rules are provided by destinations, there needs to be a way for senders to obtain these rules.

RBF leverages an extended DNS infrastructure to distribute rules. The end-hosts (or other entities on their behalf) publish rules to the infrastructure. In turn, each sender queries this infrastructure to obtain the destination's rule, similarly to the way the sender queries DNS to obtain the receiver IP address. This is the common case of distributing rules.

2.1.4 Ensuring Rules Are Authorized

To guarantee that entities participating in a rule authorize the rule, each rule is certified by a third party certifying authority, called *Rule Certification Entity*, or *RCE* for short. The RCE guarantees that all nodes whose addresses appear explicitly in a rule (*i.e.,* destinations, middleboxes and indirection routers) agree with the rule. In addition, the RCE may verify the rule for forwarding-loops before certifying it (see Section 5 and Appendices A and B). Upon receiving a packet, an RBF router verifies the rule's signature. If the verification fails, the router drops the packet; otherwise, the router applies the rule.

Since a rule cannot use a router as an indirection point without the router agreement, the ISP that owns the respective router retains full control on packet forwarding at the IP layer.

Thus, the RBF architecture consists of four main components:

- The specification of packet *rules* – their syntax, packet encoding, constraints on what rules can and cannot do. We discuss the details for the rule specification in Section 3.
- Certificate authorities called *Rule Certification Entities* (RCEs) that certify rules after checking that they are well formed, and that every destination specified in the rule agrees with (*i.e.,* has signed) the rule. We discuss the details of rule certification in Section 4.
- *Modified IP routers* that verify rule certificates and process packets as described above. We present a prototype router implementation and evaluation in Section 6.
- A *modified DNS infrastructure* that either directly resolves a host D's domain name to D's rule, or resolves D's domain name to another rule resolution server which in turn provides D's rule. We discuss the details of rule distribution in Section 4.

2.1.5 Assumptions

RBF builds on three major assumptions.

Anti-Spoofing: First, RBF assumes the existence of an anti-spoofing mechanism. This is required because rules may use source and destination IP addresses in their decision process and hence addresses must be legitimate, otherwise policy compliance cannot be enforced. Note that any solution for blocking undesired traffic inside the network requires a way to identify sources. Anti-spoofing identifies users based on their addresses. An alternative, is to identify users by their access path [108, 110], however this approach ties communications to a specific path restricting flexibility (*e.g.,* for mobility, traffic engineering, multi-path forwarding). In this dissertation we assume the use of ingress filtering as the anti-spoofing mechanism, although RBF can accommodate alternate solutions, *e.g.,* Passports [77]. The rationale behind our choice of ingress filtering is described in Section 4.

RCE Key Distribution: Second, we assume routers know the public keys of RCEs and can thus verify rule certificates. We assume the number of RCE organizations is relatively small and these keys can be statically configured at routers, akin to how browsers today are configured with the list of major certificate authorities. We discuss alternatives to this deployment in §4. Note that although we assume a small number of RCE organizations, we envisage each organization will run geographically replicated instances of their service for improved scalability and robustness.

Provisioning of Rule Distribution Infrastructure: Finally, we assume that the rule resolution infrastructure (whether DNS or the resolution servers the DNS points to) is well provisioned, akin to how major Internet services (Google, DNS, Amazon) operate today, relying on engineering approaches such as maintaining a presence at major ISPs, IP any-casting, bandwidth provisioning, and so forth. As described in §4, we make this assumption to protect against "denial of rule" attacks.

2.1.6 Summary

RBF can be succinctly described as:

1. Every packet contains a rule; there are no exceptions and no special traffic.
2. A rule is a set of forwarding directives associated to the packet by end-users; the expressivity of rules enables forwarding flexibility.
3. Each rule bears a trusted entity's signature, which guarantees the rule is authorized and safe.
4. Routers verify the rule signature and forward conforming to the rule's directives.

The described architecture supports RBF's target properties as follows:

- **Rules are mandatory:** This is easily enforced; if a router receives a packet without a rule, it simply drops the packet.
- **Rules are provably authorized:** Each rule is certified by an RCE, which guarantees that every destination or provider of router-defined function in the rule has authorized (signed) the rule. Any router and end-host can verify a rule certificate thus preventing attackers from modifying rules.
- **Rules are provably safe:** This follows from the constraints on rule syntax described earlier. We constrain rules in three important ways that enable safety: (i) an RCE certifies a rule only after checking that it cannot cause loops, (ii) a rule cannot directly modify router state and hence cannot corrupt a router, (iii) rule execution time is linear in the size of the rule. We analyze attacks using rules and defenses in §5.
- **Rules allow flexible forwarding:** We illustrate this through examples in Section 3.3.

RBF assumes routers will incorporate support for rules (*i.e.,* implement a rule forwarding layer, specialized functionality, etc.) and allows users to specify rules thereby achieving flexibility. Importantly however, RBF's design ensures that rules accommodate operator policies and concerns – for example, the path specified in a rule cannot violate ISP routing policies, providers must explicitly agree to have their routers named in rules, rules can only operate on explicitly exposed router attributes, any modification of router state or specialized packet processing can only be achieved through provider-installed functions, and so forth.

In effect, with RBF, users and network operators *share* control over how routers process packets. This feature distinguishes RBF from prior network architectures. Current IP allows users little control over how routers process their packets thereby limiting flexibility. At the other extreme: the active networks vision of routers executing code carried in packets allowed users almost unlimited control, leading to grave security concerns. In a different approach, overlay-based architectures allow users flexibility without raising new security challenges but cannot leverage in-network information and support and are thus less powerful. For example, overlay-based architectures can only drop unwanted packets at overlay nodes and hence cannot create a network that is fundamentally default-off, once the network-layer address of a node is known, it can always be attacked at the underlying network layer.

To conclude this overview, the RBF architecture consists of a data plane component that includes the rule forwarding mechanism and a control plane component that includes rule creation, certification and distribution; we describe each in detail in Sections 3 and 4 respectively.

Chapter 3

RBF Data Plane

In this section we describe the key components of the RBF data plane – how users specify packet rules and how routers verify and execute these rules.

3.1 Rule Specification

A rule can be represented by a simple transition table: based on the current value of packet attributes and router attributes, the rule may generate a new set of packet attributes and forward/drop the packet. In practice, we encode rules using an *if-then-else* tree-like structure, which has a more compact representation than a simple transition table.

Thus, a rule is represented by a sequence of actions that can be conditioned through *if-then-else* structures of the following form:

```
if (<CONDITION>)
  ACTION1
else
  ACTION2
```

Conditions are *comparison operators* applied on the packet and router attributes.
The actions can be:

1. **Forward** the packet to the underlying IP forwarding engine;
2. **Invoke** a local function available at the router;
3. **Drop** the packet.
4. **Update** the value of the packet attributes;

Routers and hosts forward packets according to the actions specified by the rule. When one of the first three actions listed above (`forward, invoke, drop`) is encountered, the

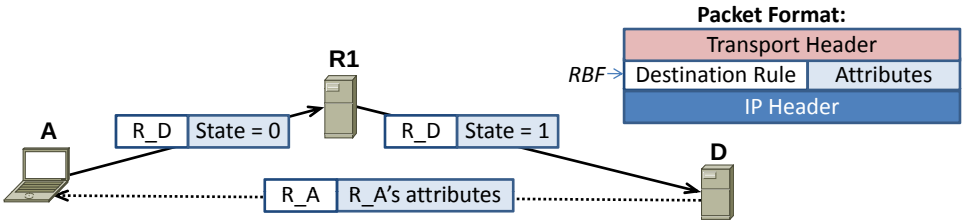


Figure 3.1: Simple indirection example.

application of the rule is ended and the appropriate action is executed. When the `update` action is encountered, the value of the packet attribute is changed accordingly, and the rule execution continues.

The packet attribute set consists of the five-tuple in the IP header (*i.e.,* IP addresses, ports, protocol type), and a number of custom attributes with *user-defined semantics*. The user-defined attributes can be used to identify different states that the packet finds itself during its journey through the network or to record information from the path. For simplicity, performance and security reasons, RBF does not allow rules to add new packet attributes. By default, user-defined packet attributes are initialized with the value zero, but can be set to a different value by senders.

Router attribute, on the other hand, can include the router's IP address, AS number, congestion level, flags indicating whether the router implements a specific functionality such as intrusion detection and so forth.

Packet attributes can be seen as having a local scope, since their semantics are associated only with the packet's rule. Router attributes, on the other hand, have a global scope, since they have semantics shared by the rules of all the packets.

Each rule has an associated *lease* that ensures the rule can only be used for a limited period of time. Routers drop packets with expired leases. For details on the lease mechanism see Section 4.3.

Also, every rule has an identifier (ID) defined as the concatenation of a hash of the rule owner's public key and an index unique to the owner, *hash(PK_owner):index*. In Section 6 we present an optimization to reduce packet overhead and identify most rules by using a hash over their content. This optimization can be used in the common case when there is no need for multiple rules with the same identifier; for example, mobile hosts may require different rules with the same identifier (see §3.3).

For example, the following rule forwards a packet to a destination `D` via a waypoint router `R1`; a packet attribute named `state` indicates whether or not the packet has already visited `R1`:

```
R_D:
  if(packet.state == 0)        //from source to R1
```

```
    if(router.address != R1)
      sendto R1
    else packet.state = 1    //at R1 sending to D
  if(packet.state == 1)
    sendto D
```

where the `sendto` action is a short form for the following: change the destination address to `D` and forward the packet using IP if `D` is not the local address, or forward it to the transport layer otherwise. Once the packet is forwarded or dropped, the rule execution stops; the packet is by default dropped if it is not explicitly forwarded by the rule. Figure 3.1 illustrates this example, when `D` is communicating with another host `A`, who's rule is `R_A`.

While RBF does not allow rules to modify the packet payload or replicate packets, RBF enables rules to `invoke` such functionality at middleboxes or routers, if available. This allows RBF to leverage recent advancements in router design that enable network operators to provide new functionality through router extensions [27,28,81]. While the router function invoked by a rule at a router may modify the packet payload, may replicate packets and may change the router state, such actions are done by code that is trusted and controlled by the network operator. The invoked router function will typically return the packet back to the forwarding layer with the same header (IP, RBF and transport), but potentially with modified payload, and multiple replicas.

The generic structure or rules enables RBF to provide a rich set of forwarding functionalities, including explicit middlebox traversal, multi-path routing, anycast, multicast, and loose source routing. We illustrate the generality of rules through a set of examples in Section 3.3.

Return rule: A packet with source S and destination D must include a *destination rule*, `R_D`, which is the rule specified and owned by D. In addition, each packet may include a *return rule*; this is the rule specified and owned by S and is used for return traffic from D to S.

Expressivity: At a high-level, our rule specification can be viewed as defining a finite state machine (FSM): the state is encoded in the packet attributes, while the input is represented by the attributes of the routers along the data path. The rule specifies the transition function of the FSM. The RBF mechanism can theoretically implement any deterministic forwarding function that can change the packet attributes only. However, due to the limited size of packets, there are forwarding functions that cannot be efficiently expressed in RBF. This limitation is similar to the impracticality of implementing complex functions with the simple FSM mechanism due to the exponential growth in the number of states. In particular, forwarding decisions based on any functions other than comparisons of packet and router attributes (*e.g.,* sum, hash, logarithm) are not practically expressible in RBF. For example, forwarding based on the sum of the addresses of routers on the path could not be easily written with rules. See Section 8.3 for a discussion on RBF's limitations.

3.2 Rule Verification

As mentioned earlier, rules are certified by a Rule Certification Entity (RCE) and all packets carry a signature that routers must verify. The verification load at routers is eased by two factors.

- First, not all routers need to verify rules but only routers at the boundaries between trusted domains. The routers inside each trusted domain simply rely on these routers to have verified the policy compliance of packets. Since trust boundary routers are typically situated at the edge of the network, they also see less traffic than core routers and can more easily accommodate the cost of verification.

- Second, routers can cache verification results by maintaining a hash of the rule and its signature. With caching, the full signature verification is only required for the first packet forwarded on a new rule (during the period in which the verification result is cached).

Thus, verifications can be limited only to border routers and, assuming a large enough cache, the verification rate is given by the arrival rate of packets with new rules. By contrast, the signature length adds to the overhead of *every* packet.

Different cryptographic solutions offer different tradeoffs between signature length, signing time (incurred only at RCEs), verification time (incurred at routers) and security. RSA is the most commonly used digital signature scheme. RSA signature verification is relatively fast compared to other signature schemes, however, RSA signatures are large. Our current RBF design assumes Elliptical Curve Cryptography (ECC) because ECC signatures (ECDSA) are shorter than RSA ones, while exhibiting similar security properties. At the same time, verification time in ECC is typically longer than RSA's. However, in practice, verification can be accelerated using ASIC-based implementations or dedicated specialized co-processors. Such implementations are already commercially available [10, 15, 16] and incorporated into network appliances and routers. Furthermore, traffic measurements [9] show that new flow arrivals represent less than 1% of the link capacity on average and less than 5% of the total number of packets, a volume that could be accommodated using commercial ECC modules [10,15] or recent research proposals [70,112]. We evaluate the size overhead of ECC signatures compared to RSA as well as the verification overhead of RSA in Section 6.

Note that the RBF design is independent of the signature scheme used. If fast enough supporting hardware is built, other signature schemes such as DSA or very short signature schemes such as those proposed in [40, 84] could be used. Since RSA is the only signature scheme used at scale today, there are few high speed commercial products for other signature types. However, this may change in the future.

3.3 Usage Examples

RBF gives end-users four basic types of control:

1. *Block* unwanted packets in the network;
2. Redirect packets through a sequence of *waypoints*;
3. Use *enhanced functions* at routers (if available) and *middleboxes*;
4. Use *router state* in the forwarding decision and record such state (*e.g.,* the maximum congestion level).

Next, we present several examples to illustrate RBF's flexibility.

Port-based filtering

A web server, `D`, can use the following simple rule – registered under its DNS name – to make sure that it receives only packets destined to port 80:

```
R_filter_port:
  if(packet.dst_port != 80)
    drop;
  sendto D
```

Middlebox Support

In addition to accepting traffic directly on port 80, `D` can use the following rule to route all the other incoming traffic through a packet scrubber [7, 13], deployed either by `D`'s provider, or a third party:

```
R_mbox_port:
  if(packet.dst_port == 80)
    sendto D
  else
    if(packet.state == 0)          //before scrubber
      if(router.address != Scrb)
        sendto Scrb
      else
        packet.state = 1           //mark scrubbed
        invoke Scrb_service
    else if(packet.state == 1)
      sendto D                     //packet has been scrubbed
```

Thus, similar to other previous proposals [98, 104], RBF provides explicit support for middleboxes, such as WAN optimizers, proxies, caches, compression or encryption engines, transcoders, SSL offloaders, intrusion detection (IDS) boxes.

Secure Middlebox Traversal

In the previous example, an attacker can directly send a packet with the attribute values set such as to appear that the packet has already visited the middlebox. The destination can protect against this behavior in two ways, which we describe next.

In the first approach, `D` can simply ensure that the packet does indeed arrive from the middlebox when the `state` attribute is set to the value of one (three lines have been added to the previous rule):

```
R_mbox_port_antispoof:
  if(packet.dst_port == 80)
    sendto D
  else
    if(packet.state == 0)          //before scrubber
      if(router.address != Scrb)
        sendto Scrb
      else
        packet.source = Scrb       //NEW -- mark from Scrb
        packet.state = 1           //mark scrubbed
        invoke Scrb_service
    else if(packet.state == 1)
      if(packet.source != Scrb)    //NEW -- check if from Scrb
        drop                       //NEW -- drop if not
      sendto D                     //packet has been scrubbed
```

This rule relies on the anti-spoofing mechanism used in RBF to block packets with spoofed source address attributes.

Note that to avoid legitimate packets being dropped by the anti-spoofing mechanism, the source address attribute should be set at all off-path waypoints and routers that change the destination address; we omit this in the presented rules for readability purposes.

In a second approach, the destination can use stronger cryptographic guarantees. For this purpose, the middlebox must implement a certifying function that creates a cryptographic proof to certify that the packet visitied it. The destination must implement another function that verifies these proofs before delivering the packet to the application. These functions are invoked by the rule:

```
R_mbox_port_crypto:
  if(packet.dst_port == 80)
    sendto D
```

```
else
  if(packet.state == 0)          //before scrubber
    if(router.address != Scrb)
      sendto Scrb
    else
      packet.state = 1           //mark as scrubbed
      invoke Scrb_service
  else if(packet.state == 1)
    packet.state = 2
    invoke Certify               //create proof at Scrb
  else if(packet.state == 2)
    if(router.address != D)
      sendto D
    else
      invoke VerifyAndDeliver    //verify at D
```

One example for the `Certify` function at the middlebox signs an attribute that initially represents a hash of the packet header (or payload) content that does not change during the packet forwarding. The `VerifyAndDeliver` function checks if the certification function was actually invoked at the middlebox; for this purpose it, has to know the public key of the middlebox. If the verification condition is met, this functionality delivers the packet to the transport layer. This process can be generalized to arbitrary rules, middleboxes signing this attribute one after the other. To determine whether a sequence of signatures is valid, the verification function must apply static analysis and check whether the packet attributes can correspond to the sequence of signatures. Note that many other cryptographic procedures can be used in this context, *e.g.,* secret hash functions, secret keys, *etc.*.

Also note that to increase the performance, the certifying function can be applied to an entire batch of packets instead of individual packets. In this case, the entire batch must be buffered and checked at the receiver before being delivered to the application.

DoS Protection

To protect against DDoS attacks, a server, `D` can create a custom rule for each client, which drops packets from any source other than the client. By controlling the number of rules active at a given time, `D` controls the maximum number of active clients (each rule has associated a lease period). An example of a rule similar to a network capability [108, 110] is:

```
R_filter_src:
  if(packet.source != requester_IP)
    drop;
  ...//rest of the rule
```

Similarly to capability based architectures [108,110], our solution is based on the premise that destinations are able to grant rules on demand, and that any requester can ask for a destination's rule. In RBF, this task falls to the rule resolution infrastructure and raises the possibility of a "denial of rule" attack on this infrastructure (akin to denial-of-capability attacks in capability-based systems [87]). We present the details of rule resolution and discuss denial-of-rule attacks in §4.

To see how different types of uses can be naturally combined in RBF, note that this preamble can be applied to any other rule. For example, one can create a rule that only accepts packets from one source on one port, or a rule that uses a middlebox but only accepts packets from a given address prefix, *etc.*

DTN Support

An ISP can partner with a third-party storage provider `gooInc` to offer a delay-tolerant network (DTN) service [52]. The ISP uses an intermediate router `Ind` that tracks the availability of a disconnection-prone destination `D`; this availability is indicated by the `D_online` router attribute. Packets destined to `D` are routed to `Ind`. When `D` is reachable, `Ind` forwards the packets directly to `D`; otherwise it forwards `D`'s packets to a DTN storage box `St` provided by `gooInc`. The following rule captures this functionality:

```
R_DTN:
  if(packet.state == 0)                  //to Ind
    if(router.address != Ind)
      sendto Ind
    else
      if(router.D_online == TRUE)
        packet.state = 1                 //to host
      else
        packet.state = 2                 //to DTN-box
  if(packet.state == 1)
    sendto D
  if(packet.state == 2)
    sendto St
```

In this example we have assumed that the DTN host returns online having the same IP address as earlier (`D`). Otherwise, the DTN host must update its rule in the rule distribution infrastructure every time it gets online to reflect the new address, but can still recover the packets sent to it while offline from the storage box `St`.

Mobility

The mobile node `M` changes its network IP address due to physical movement. In RBF, `M` can continue an existing communication without having to re-establish it. To achieve this,

M creates a rule for the new address with the same identifier as the rule used in the existing communication, and places it in the packet as the return rule.

Multicast

For security reasons, RBF does not support packet replication, and thus multicast cannot be implemented entirely at the RBF layer. Instead, multicast can be implemented by invoking multicast functions deployed by ISPs at some of their routers; these functions maintain (soft) state at routers to create a multicast (reverse path) tree. This approach implements essentially an overlay multicast solution, which leverages the IP multicast functionality at on-path routers. For simplicity, here we consider only single-source multicast trees.

Consider source S that wants to send packets to a multicast group uniquely identified by G. S advertises (*e.g.,* on the web) that receivers should use the following registration rule:

```
R_multicast_registration_G:
  if(router.multicast_available and
                    packet.crt_router != router.address)
    packet.crt_router = router.address
    invoke multicast_registration
  sendto S
```

where the crt_router attribute makes sure multicast registration is called just once at each multicast router.

When joining the multicast tree, a receiver, D, sends a registration packet using the above rule. Prior to sending this registration packet, D creates a rule to receive the multicast packets sent by S and inserts it in the packet's payload:

```
R_mcast_forwarding_to_D:
  packet.dst_port = PORT_D_LISTENS_MCAST_G
  sendto D
```

The packet payload also contains the identifier G. The first multicast enabled router R processing the registration packet, stores the mapping G→R_mcast_forwarding_to_D. R creates its own rule to receive these multicast packets, replaces D's rule in the packet, and sends the packet further. The registration continues recursively until it reaches a multicast-capable router on the (reverse) path from D to S that already stores an entry for group G.

To send a multicast packet, S sends a copy of the packet to every router from which it has received a registration.

On top of the vanilla multicast functionality, this approach can easily implement other functionalities, such as access control and traffic accounting, which have been previously proposed to "fix" the IP multicast [67].

Path Selection

We consider a web server D that wants traffic from its clients to travel along low congestion paths. Similar to prior proposals [36], D achieves this using a form of loose source routing although in this case routes are selected by the destination. We assume D has knowledge of some small number of potential indirection points located in different ASes. While in general these can be routers that D discovers through agreements with ISPs or middleboxes offered by third-party service providers, our example below assumes these are routers.

Initially, D simply uses normal routing and records the congestion level on all packets it receives using the following rule:

```
R_congestion:
  if(router.congested >= packet.congestion_level)
    packet.congestion_level = router.congested
  sendto D
```

If D notices that incoming traffic from a source A is seeing high congestion, then D asks A to switch to a new destination rule R_congestion_lsr1 that routes traffic from A to D via a series of indirection points that D selects based on its knowledge of AS topology. D effects this switch by simply setting R_congestion_lsr1 as the return rule on packets from D to A. An example of such a rule using two indirection routers IR1 and IR2 is as follows:

```
R_congestion_lsr1:
  //record congestion
  if(router.congested >= packet.congestion_level)
    packet.congestion_level = router.congested

  //follow the loose path
  if(packet.state == 0)          //to IR1
    if(router.address != IR1)
      sendto IR1
    else
      packet.state = 1
  if(packet.state == 1)          //to IR2
    if(router.address != IR2)
      sendto IR2
    else
      packet.state = 2
  if(pkt.state == 2)             //to D
    sendto D
```

Notice that D continues to measure the congestion level on all incoming traffic and can hence continue experimenting with alternate incoming routes until it finds a satisfactory one.

Also, note that unlike IP source routing and similar to other proposals [86, 90], RBF's loose paths are "secure", bearing the approval of the waypoints; users cannot arbitrarily select loose paths but these have to be contracted with the ISPs.

Multiple Paths

As we have shown in the previous example, a rule can encode a loosely specified path. This enables RBF to route packets along multiple paths. For example, a multi-homed host `D` can choose its network provider similar to [86, 90]; *e.g.,* `D` could select one provider for VoIP traffic and another provider for large data transfers. `D` can also use multiple paths for the same flow to increase throughput. For this purpose, `D` could use a rule that encodes two distinct paths, and ask the sender to set one attribute with random values to select between the two paths; this selection could also be done through a random attribute at a router. In another example, `D` could use two rules with the same name that encode different loosely specified paths and alternate the two rules in the return packets.

On-path router function – Caching

Consider an ISP `I` that deploys caching functionality at some of its (border) routers. A web-service `D` can contract with `I` and use this functionality. For this purpose, `D` creates and publishes the following rule:

```
R_caching:
  if(router.caching_available and
          packet.crt_router != router.address)
    packet.crt_router = router.address
    invoke caching
  sendto D
```

where the `crt_router` attribute makes sure the caching functionality is called just once at each caching-enhanced router.

In this example, the caching functionality can decide to respond to the requester directly and not forward the packets further to `D`, which reduces the latency for the requester and the traffic at `D`.

Anycast and on-path IDS function

In this scenario, we consider a CDN `akInc` that wants to direct clients to close by servers. We assume that `loc` is a well-known packet attribute in which a source may optionally record its AS number, and the CDN provider redirects packets based on their `loc` attributes and its knowledge of the global AS-level topology. The CDN provider `akInc` thus publishes a simple rule `R_ak` of the form:

```
R_ak:
  if(packet.loc == L1) sendto D1
  else if(packet.loc == L2) sendto D2
  ...
  else sendto Dn                    //default (undefined location)
```

A source `A` then sets the `loc` attribute in its outgoing packets and uses `R_ak` as its destination rule.

Now let's say that `akInc` is interested in leveraging the IDS infrastructure deployed by a major ISP. On entering an agreement with the ISP, `akInc` is informed of the requisite router attribute `ids_avail` and handle `ids_func` by which these IDS services might be invoked. Then, `akInc` updates its rule as follows:

```
R_ak_ids:
  if((packet.used_ids == FALSE) and
                  (router.ids_avail == TRUE))
    packet.used_ids = TRUE
    invoke ids_func
  if(packet.loc == L1) sendto D1
  ...
```

This allows `akInc` to quickly incorporate IDS functionality into its deployment and allows the provider to offer use of its IDS service without revealing details of its IDS deployment such as the IP addresses of IDS routers.

Note that this form of anycast may sometimes lead to large rules (*e.g.,* the number of servers can be large, the selection criteria can be complex). RBF users can augment the above type of implementing anycast by using specific in-network anycast functionality deployed by ISPs. Such functionality can use arbitrary metrics (*e.g.,* a network map) to select the servers to which to send the packet. However, when invoked, the anycast router function needs to tunel the packet to the intended anycast destination since the rule inserted by the sender is not authorized to send packets to the (unknown by the rule) anycast destination.

Network Probing

With RBF, end-users can record network monitoring information from the path, *e.g.,* some of the router attributes can be elements of the management information base (MIB) table. For example a rule can record from the path the maximum/minimum link bandwidth, forwarding table size, number of incoming/error packets, number of neighbors, queue size, *etc.* Also, users can verify whether the path passes through a certain link type or through a particular AS/country. Furthermore, specialized router probing functions, such as proposed in [96], can be invoked for many other use cases.

Path Identifier

A rule that records the last four routers of the path can easily be created with RBF; this is similar to Pi [107]. Such a rule can be used to detect attackers, in case a partial anti-spoofing mechanism is used and some of the hosts can spoof their addresses. Further potential refinements of the path identifier recording (such as to record only the routers for which the destination address was not learned through IGP, see [107]) can be done by invoking specialized router functions.

On-path router function – Energy Saving

ISPs can deploy a router forwarding mechanism that saves energy at routers at the expense of minor extra delay experienced by users [85]. Users are informed by the requisite router attribute `green_avail` and handle `green_func` by which this service might be invoked. Energy cautious users explicitly invoke this functionality to reduce the global energy usage:

```
R_green:
  if(router.green_avail == TRUE and
              packet.crt_router != router.address)
    packet.crt_router = router.address
    invoke green_func
  sendto D1
```

On-path router function – Content-based Routing

Recently, several research proposals [46, 69, 76] argue for embedding in routers an anycast primitive that forwards packets to the closest copy of the data based on the name of the requested data. This would help to reduce the cost of the anycast service, increase data's availability and reduce latency (as DNS queries might be be avoided). RBF can implement content-based routing by invoking route-by-name functionality available at enhanced routers:

```
R_route-by-name:
  if(router.route-by-name_available and
              packet.crt_router != router.address)
    packet.crt_router = router.address
    invoke route-by-name
  sendto DefaultDst
```

where `DefaultDst` is the original holder of the data. The `route-by-name` functionality reads the data name from the packet (*e.g.,* this could be applied to HTTP packets), and if it has an entry for this data, it replaces the current rule with the rule of the next hop towards the closest data copy, and then forwards the packet. The original destination rule used by the sender should also be kept in the packet and be returned to the sender as the return rule.

3.4 Source Rules

In some cases, the source may also desire control over how its outgoing packets are forwarded. For example, the source might want to send packets through an anonymizer, behavior that should not be under the control of the destination.

For this purpose, we incorporate the ability to have a *source rule* in packets in addition to the destination rule. A packet is always forwarded first on the source rule (if present) and once the source rule has been completed, the packet is forwarded as per the destination rule. We use a special (one bit) packet attribute to denote which rule is currently active and only allow this attribute to be set, thus ensuring that control does not return to the source rule once the destination rule has been activated. This can be enforced through runtime checks made by the rule execution engine (*i.e.,* by specification this operation is not allowed) or can easily be verified by static analysis before rule certification.

For example, while the scenario discussed in the previous section used loosely specified paths and congestion recording based on destination rules, one could achieve a similar effect using source rules for outgoing traffic, and the selected paths can be different paths to/from the Internet core.

Note that the same result as when using source rules can be achieved by simply encapsulating packets. Specifically, the host `S` could encapsulate packets sent to host `D` with an outer header which contains a rule desired by `S`; the recipient of that rule simply removes the outer header and continues to forward the packets using `D`'s rule. However, having support for source rules built in RBF does not require middleboxes/routers to be aware of encapsulation nor to strip packets of the source-added headers.

Importantly, note that not all (destination) rules are compatible with the use of source rules. Specifically, only the destination rules that accept traffic from any source address can be used with arbitrary source rules. For example, in the communication between hosts `S` and `D`, `S` can use a source rule that sends packets through an indirection host `M`, only if `D`'s rule accepts packets generated by `M`. Otherwise, `D`'s rule will simply drop packets forwarded through `M` since, in general, the anti-spoofing mechanism prevents `M` from sending packets with `S`'s source address. These conflicts can simply be discovered by marking some of the rules as *incompatible* with the use of source rules. Specifically, if the destination's rule is incompatible with source rules, the sender must not use any source rule (or otherwise it cannot communicate with the destination). It is easy to statically verify the property of being incompatible with source rules. For simplicity, this property could also be signaled through a bit in the rule identifier.

One could generalize source rules to having a stack of rules in the packet, which are executed one after the other. We leave to future work the decision of whether to incorporate this abstraction within the rule mechanism.

Throughout this dissertation, unless otherwise specified, we assume there is no source rule present in packets.

3.5 Functions on Both Forward and Reverse Paths

Some middleboxes and router-defined functions change the data payload of packets before sending them to the destination. Examples of such functionalities are compression engines, encryption engines, transcoders and TCP accelerators. Examples of functionalities that are not in this category are firewalls, intrusion detection systems, load balancers, deep packet inspection boxes/packet classifiers, packet markers, simple proxies and some redundancy elimination boxes.

Assume host S communicates with host D and that the rule of host D uses such a box M (middlebox or router) that implements a function modifying the data payload. If the communication between S and D requires end-to-end reliability (usually the case) D has to request retransmissions of lost packets from M rather than the original source of packets S, since the bytes received by D do not correspond to those sent by S.[1] To enable D to request retransmissions from M, the traffic returned from D to S should also pass through M. In other words, instead of having only the direct traffic from the source to the destination going through M (as specified in the destination's rule), the return traffic from the destination to the source should also pass through that middlebox or router. For example, a common scenario in today's Internet is for such types of functions to terminate TCP connections, *i.e.,* have two TCP connections one between S and M and one between M and D.

In addition, even functionalities that do not change the payload content of packets might benefit from being on the return path as well, *e.g.,* IDS may be able to more easily reconstruct the state of some protocols, caches need to know the state of the TCP stack to be able to take over the response from destinations, *etc.*

We call these types of middleboxes and router-defined functions that require to be on both the forward as well as the return paths as "*bidirectional*", in short BD.

Therefore, the reverse traffic from destination D must pass through the BD box M (middlebox or router) on its return to source S. Note that M is actually specified in D's rule and thus, D knows about M's existence. We discuss two separate contexts: (i) M is a middlebox specified in D's rule and (ii) M is a router that implements a function invoked by D's rule. In the latter case D does not know M's identity. We assume R_D is D's rule (containing M), R_S is S's rule.

i) Middleboxes on return path: We describe two methods by which the reverse traffic from the destination D can pass through the middlebox M on its way to the source S.

1. *Explicit:* RBF enables the explicit placing of middleboxes through rules. Therefore, in order for M to be situated on the path from D to S, the natural solution is to have M appear in S's rule, R_S, as well. In this way, M is explicitly located on both directions of the communication and S is fully aware of M's (off-path) presence for the traffic from

[1] Applications could require retransmissions of entire application data units, but this may lead to poor throughput.

D. This approach does, however, require S to communicate with M on the control plane and, most likely, create a rule for sole purpose of communicating with D.

2. *Source rule:* D can use the source rule R_M which directs packets to M; from M packets are forwarded using R_S. In this way, S does not need to create any new rule if simply does not care about the middleboxes used by D. D and M can create rule R_M out of band, *e.g.,* at the same time with R_D. Alternatively, M could place R_M in the payload of some of the packets on their way from S to D; however, this latter approach would not scale to multiple middleboxes.

The aforementioned limitation of the ability to use source rules applies when source rules are used to place BD middleboxes on the reverse path. Specifically, in our previous example, R_S should accept packets generated by M, and cannot drop all packets not initiated by D, since, in general, the anti-spoofing mechanism may prevent M from sending packets with D's source address. However, as we will discuss in Section 4.2.1, this limitation can be addressed by having endpoints (or their RCEs) create DoS protection rules that also accept packets from the middleboxes used in the endpoints' rules, in addition to the endpoints.

ii) Router-defined functions on return path: We now discuss how router Rtr implementing a BD function invoked by D's rule R_D can be placed on the return path from D to S.

1. *Implicit:* Since routing paths are not controlled by end-users, ISPs should typically deploy BD functions only at routers close to hosts, which are guaranteed to be on both the forward and the return path of packets. In this case, the BD router function invoked by R_D and located on Rtr can simply inform the forwarding module of Rtr to capture the return packets associated with the respective flow and forward them through the BD function.

2. *Source rule:* The BD function could also be implemented at a router in the network core. However, in this case, if the path changes, the communication will be broken and needs to be re-established. To ensure packets go through Rtr on their way back from D to S, source rules can be used as described above for the case of BD middleboxes. Since Rtr is not known by D before S communicates with it, Rtr's rule must be communicated to D at runtime, *e.g.,* by including it in some of the packets sent to D.

Importantly, note that rules using a BD function must not dynamically select waypoints on the path based on router attributes or, otherwise, not all packets may pass through the BD box (such a rule would be a poorly designed rule).

Chapter 4

RBF Control Plane

In this section, we describe the RBF mechanisms for rule creation and certification (§4.1), rule distribution (§4.2), lease enforcement (§4.3) and anti-spoofing (§4.4).

Setup: In RBF, ISPs must provide their clients rules to access a local DNS server, as well as a Rule Certification Entity (RCE), which can certify clients' rules. This information can be provided through a modified DHCP service, similar to the way ISPs or organizations provide the IP address of DNS servers today.

4.1 Rule Creation and Certification

To receive traffic, a client must create a rule that allows one or more sources to send traffic to it. Before distributing this rule, the client must ask an RCE to certify it. If a rule is not certified, the RBF routers will drop all packets using that rule.

RCE certification guarantees that rules obey the policies of all stakeholders. In particular, certification guarantees the following properties:

1. Every destination in the rule (*i.e.,* any address that appears as an argument of a `sendto` instruction) has agreed to receive packets using that rule;

2. The operators providing router functions invoked by the rule approve the rule behavior (*i.e.,* the rule invoking those functions);

3. The rule cannot cause infinite loops;

4. The rule cannot bypass ISP routing policies (see §5).

A client can either create rules itself and directly ask an RCE to certify these rules, or use a trusted DHCP-like service to create and certify rules on its behalf. In the remainder of this section we present the former case. We simply note here that through DHCP-like

services (the latter case) ISPs can use rules offer a broader ranger of services to clients, while clients are not aware of the use of rules.

As mentioned earlier, the ISP provides each client with a rule to access an RCE that has a contract with the ISP. The following example shows a possible rule that allows a client D to access an RCE named C:

$$\mathrm{R}_{D\rightarrow C}\colon \ \mathrm{if}\ \ (\mathrm{source} == \mathrm{D})\ \ \mathrm{sendto}\ \ \mathrm{C}$$

Before certifying a rule, an RCE verifies that the rule has been authorized by each destination that appears in the rule. The client who has created the rule authorizes it by simply signing the rule with its private key. A client that appears in the rule as a destination, other than the rule's creator, will first verify that the content of the rule obeys its policies before signing the rule. For example, an intrusion detection box may verify that the destination indeed belongs to a client allowed to use the service (*e.g.,* based on a contract between the client and the provider of the intrusion detection service), a waypoint router may verify that the final destination is allowed to use source-routing, *etc.*

Let $(K_D,\ K_D^{-1})$ denote the (public, private) key pair of client D, and let $\mathtt{IP}_D$ be the IP address of D. To prove to an RCE that the client signing the rule with private key K_D^{-1} indeed owns IP address $\mathtt{IP}_D$, client D sends a certificate along with the signed rule that binds its public key K_D and IP address $\mathtt{IP}_D$. This certificate is signed by an entity T, *i.e.,* $[IP_D, K_D]_{K_T^{-1}}$, where K_T^{-1} represents the private key of T. Clearly, the RCE must trust entity T. In fact, in our solution we will assume that T is itself an RCE.

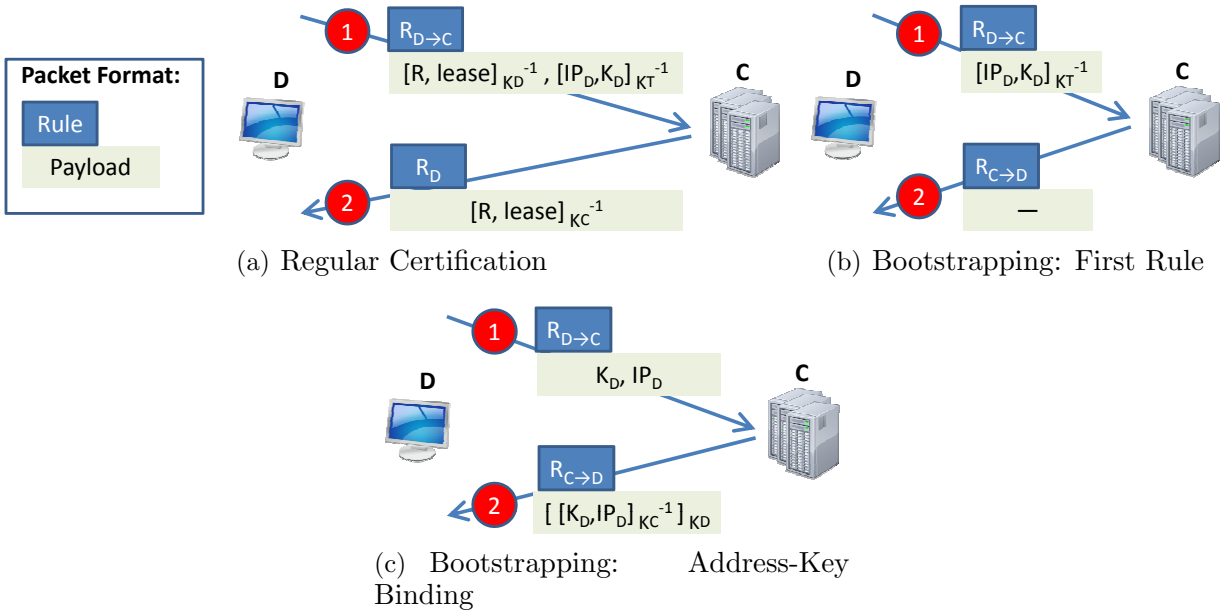
Next, we present the rule certification process in detail, initially for the case in which the rule has a single destination, and then for the case in which the rule has multiple destinations or waypoints/middleboxes.

Certify single-destination rules: Assume destination D wishes to certify a rule R that forwards packets only to its address $\mathtt{IP}_D$, *e.g.,* `R: sendto` $\mathtt{IP}_D$. Also, assume D already has a rule $\mathtt{R}_D$ on which it can be reached by the RCE C. D obtains this rule as part of the bootstrapping process, which we discuss later.

Fig. 4.1(a) shows the certification of D's rule, R, by C:

1. Host D signs rule R with its private key, and sends it to C using rule $\mathtt{R}_{D\rightarrow C}$. In addition, D sends the certificate binding its public key and address, *i.e.,* $[IP_D, K_D]_{K_T^{-1}}$. Upon receiving this request, C verifies the certificate as well as the signature of the requested rule. These ensure that the request has been made by the owner of $\mathtt{K}_D$ and that the requester is also the owner of $\mathtt{IP}_D$. In addition, C verifies that R is well formed (see §5). For C to learn $\mathtt{K}_D$, $\mathtt{K}_D$ should also typically be contained in the same packet or in an earlier one (or instead of using a signed binding between the public key and the IP address, one can use an encrypted such binding).

2. If rule verification succeeds, C signs the rule with its private key and sends it back to D using the return rule in its certification request, $\mathtt{R}_D$. At this point, host D can



(a) Regular Certification

(b) Bootstrapping: First Rule

(c) Bootstrapping: Address-Key Binding

Figure 4.1: Rule Certification

distribute rule `R` to other hosts directly (as a return rule) or through DNS.

The certification procedure (Fig. 4.1(a)) needs only to guarantee the authenticity of the request. Since rules are public, confidentiality is not a concern. Since the lease is an absolute value (§4.3), the only effect of replaying rule requests is increased traffic at the RCE. The maximum lease value that `C` can sign for a rule is negotiated between `D`'s ISP and `C`. Furthermore, RCEs can limit the number of clients contacting them and can limit each user's certification rate, as we discuss in this section.

Timeouts should be used to handle lost packets. We discuss the load of RCEs later this section.

Certify multiple destination rules: In this case, every destination (*i.e.,* any host, middlebox, or waypoint router that appears as an argument of a `sendto` instruction) in a rule must agree to receive packets on that rule, *i.e.,* the rule must respect its policies. In particular, every such destination must sign the rule. One of the destinations, `D`, collects the signatures of all the other destinations along with their certificates binding their public keys to their addresses. `D` then sends this information to its RCE. In turn, the RCE verifies that all destinations in the rule have signed the rule and sends the signed rule back to `D`. The lease signed by the RCE has the minimum duration between the requested lease and the leases of all the certificates binding the addresses and the keys of the participants.

Certify rules invoking functions: Operators providing router functions can restrict which rules can invoke these functions. The certification process is similar to certifying multiple destination rules. The identifiers of functions whose invocation requires authorization are represented as hashes of public keys. RCEs certify a rule containing such an invocation only if the rule is signed with the private key corresponding to the function identifier.

Bootstrapping: To certify rules, client D must contact an RCE. Recall that the ISP provides D with a rule to access an RCE, C. In addition, D needs to provide C with a return rule. Otherwise, C cannot send the certified rule back to D, since RBF requires every packet to have a rule. The question is how can D get its *first* rule certified by C, without having a return rule in the first place?

To get around this challenge, D asks C first to certify a rule that allows C to send packets back to D, *e.g.,*

```
R_{C→D}:
  if (packet.source == C)
    sendto D
```

Upon receiving such certification request, C certifies it and sends the certificate back using the certified rule itself, $\mathtt{R}_{C\to D}$. If the certification request has no return rule, C simply sends the certificate back using the certified rule. Fig.4.1(b) illustrates the first rule's certification.

Recall that when sending a rule certification request, client D also needs to provide the RCE C with a certificate issued by a trusted authority, T, which signs the binding between D's key K_D and its address IP_D. In general T can be any authority trusted by C. For example, T can be the ISP that provides service to D. For simplicity, in this dissertation, we assume that T is also an RCE and, in particular, that T and C are the same.

Fig.4.1(c) shows how client D obtains a certificate of the binding between its address and public key. First, D sends to C a request containing only K_D and its address IP_D with no return rule (at this point D has no rule). Upon receiving this request, C signs the binding between K_D and IP_D, and encrypts it with K_D to guarantee D indeed owns the private key K_D^{-1}. To return this packet to D, C creates a rule with the same content as $\mathtt{R}_{C\to D}$, but unlike $\mathtt{R}_{C\to D}$, this rule is named using one of C's keys and not D's key.[1] The lease of the binding also represents a contractual agreement between the D's ISP and the RCE C.

This procedure for obtaining the binding between a public key and an address, relies on the fact that D is indeed the actual owner of IP_D. A malicious router on the reverse path from C to D or another host[2] that can spoof D's IP and can intercept its incoming packets could make such requests for IP_D. These attackers can forge rules named with their keys that send traffic to victim D's IP. We discuss this attack in more detail in §5.

[1]Note that in general, an RCE can always contact a host by just creating a rule to its address.

[2]E.g. a host on the same wireless network as the victim.

Source Rules: We note that there is no difference in the certification procedure between rules logically imposed by sources and those of destinations.

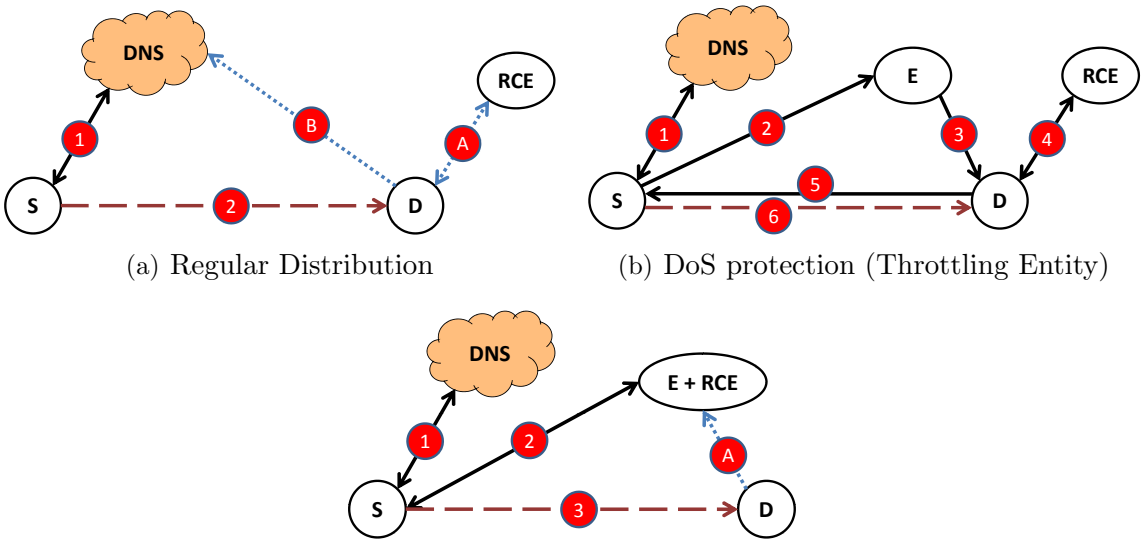
RCE load and availability: To control its certification load, an RCE can rate-limit the number of certification requests that it processes from each individual client. Clients are identified by IP address; the anti-spoofing mechanism prevents clients from impersonating each other. Alternatively, clients can be identified by the "personalized" rules provided by the ISP to each customer to access the RCE; such rules may have a finer granularity than the anti-spoofing mechanism. RCEs can indirectly protect themselves against link-level DoS attacks by controlling the number of clients under contract.

RCEs must be highly available to enable rule certification at any time. RCEs can meet this requirement by using multiple servers and multiple sites. ISPs and destinations can protect themselves against RCE unavailability by contracting with multiple RCEs.

Automatic rule agreement: A host or middlebox appearing in a rule with multiple destinations needs an automatic way to check if it agrees with the rule, in the case when the rule was created by one of the other destinations. We expect an out-of-band process by which operators of middleboxes decide which sources/destinations they are willing to act for and the nature of acceptable rules (lease durations, other middleboxes, etc.). Given this predetermined information, a rule participant `M` can automatically check if it agrees with a rule created by another rule participant by verifying the following properties: (1) only sources agreed by `M` are allowed to send packets on the rule, (2) the rule lease does not exceed `M`'s desired duration for the rule. Optionally, some participants may verify whether certain destinations are reached after themselves; *e.g.,* middleboxes may want to verify whether packets are indeed going to the desired destination afterwards. These properties are verified by statically analyzing the rule. We leave implementing such tools to future work (see §10).

RCE Key Distribution and Revocation: In this dissertation we do not explore solutions for the distribution and revocation of RCE keys to routers. Here, we simply mention two possible approaches towards this goal. In one approach, RCE keys could be distributed and revoked using DNSSEC. For example, in the `txt` or other RR type, one DNS entry contains the number of RCEs and, for each RCE, there is one DNS entry (based on its index such as "ID24.rce") that contains the RCE's key. Routers periodically update the RCE keys. In another approach, RCEs could be deployed along AS boundaries, such that each AS would have its own RCE. This approach has the advantage that additional security can be enforced, *e.g.,* the trust in some RCEs can be restricted to their own address ranges. Secure BGP can be used to distribute RCE keys in this case, but at the expense of extra complexity.[3]

[3]Note that DNSSEC could also potentially be used to distribute keys when RCEs are deployed along AS boundaries.



(a) Regular Distribution

(b) DoS protection (Throttling Entity)

(c) DoS protection (Throttling Entity + RCE)

Figure 4.2: Rule Distribution
(solid lines = rule lookup process; dashed lines = data communication; dotted lines = setup)

4.2 Rule Distribution

RBF uses an extended DNS infrastructure to distribute rules, as illustrated in Fig. 4.2(a). The destination `D` creates and certifies a rule for itself (step A) and inserts it into the DNS (step B). A sender `S` that wants to contact `D` looks up `D`'s name in the DNS; the DNS is extended to return `D`'s rule rather than its address (step 1).[4] After obtaining a rule to `D`, `S` directly sends packets to `D` (step 2).

For practical purposes, the rules of the DNS root servers should have long leases or never expire. In this way, tedious reconfigurations or refresh protocols are avoided, similar to how today DNS root servers have long-lived addresses.

In Section 3.3 we pointed out that rules can be used to block DDoS attacks. This relies on (1) the ability to distribute customized rules to different senders (*i.e.,* give a sender `S` a rule that drops all packets not generated by `S`) and on (2) the ability to protect the rule distribution itself from DoS attacks. We discuss distribution mechanisms to achieve these next.

[4]For example, DNS could be incrementally extended with a new type of record to serve rules rather than addresses.

4.2.1 Distributing Rules for DDoS Protection

To protect against DDoS attacks, client D can contract with a large entity E, and redirect its DNS entry to E, by registering E's rule under its DNS name. Fig. 4.2(b) illustrates this approach. DNS will reply to a lookup for D's name with E's rule (step 1). The DNS entry that contains E's rule must belong to a new type of DNS RR. This new class of entries is returned directly to clients by DNS resolvers. Upon a receipt of such an answer to its DNS query, the requester will continue the DNS lookup by contacting E (step 2). E rate-limits rule requests and forwards them to D (step 3), thus protecting D from DoS attacks. For the authorized requesters, D creates rules (step 4) and replies back to the requesters (step 5). E forwards requests to D conforming to a policy (see §3.3), which can be updated by D at any time.

Note that some malicious users may still get their requests forwarded by E and authorized by D. To alleviate this attack, E can employ fair queuing across senders, and D can blacklist known attackers at E. Such an approach offers a protection similar to network capabilities that apply per-source fair queuing at routers [78].

Fig. 4.2(c) presents an alternative solution, in which E directly creates rules for requesters. In this case, E has to incorporate the functionality of an RCE. To enable E to generate rules on its behalf, D provides it a rule template parameterized by the requester's address (step A). For example, the template could be similar to the rule for protecting against DoS presented in §3.3: `if (packet.source != <requester>) drop; else ...`; E uses this template to generate a unique rule per client. Upon receiving a lookup request from client S, E replaces the `<requester>` template parameter with S's address(es) extracted from S's return rule, and sends the newly created rule to S (step 2). In addition to the rule template, D informs E of a rule granting policy, which also contains the maximum number of active rules E can grant at a given time. If under attack or heavy load, D informs E to stop granting new rules, to reduce the maximum number of rules it can grant, or to block specific clients.

Compared to the approach presented in Fig.4.2(b), the solution from Fig.4.2(c) avoids involving the destination in the rule certification process; however, it requires some of the RCEs to be able to withstand DoS attacks (those that provide this service) and may require more frequent policy updates between D and E.

Typically, to avoid the limitations of using middleboxes that should be on both the forward and reverse paths (see §8.3), the rules created to protect against DDoS attacks should allow packets generated from all the addresses expressed in the requester's rule.

4.3 Rule Leases

The lease is an expiration time stamp certified along with the rule description. A router drops a packet if its current time exceeds the rule expiration time.

To implement leases, we propose that all routers and RCEs are synchronized, via NTP

[26] as recommended by router manufacturers [32]. In this context, a lease is an absolute expiration time value (see Section 6.1 for details on our implementation).

In the absence of global clock synchronization, we next describe an alternative solution in which lease times are relative to the RCE.

4.3.1 Alternative Lease Mechanism Not Using Global Synchronization

At a high level, each router maintains one timer per RCE and the current time at routers is set and updated from the flowing packets. For this purpose, rules also have associated a creation time besides the lease expiration time.

Each RCE associates to each certified rule two time stamps: a certification time `cer_time`, and an expiration time `exp_time` (when using NTP only the expiration is required). The timestamps are *relative* to the time of the RCE certifying the rule.

Each router maintains a clock t_C for each RCE `C`. t_C advances at the router's internal clock rate. t_C is initially in an uninitialized state and is initialized to the `cer_time` value of the first forwarded packet that is certified by `C`.

Upon receiving a packet certified by `C` with certification time `cer_time`, the router updates t_C as follows:

$$t_C = \max(t_C, cer\_time) \tag{4.1}$$

In other words, the current time estimate for RCE `C` is set to the most accurate value; if the `cer_time` is higher than t_C it means that the current time estimate is behind the actual time of RCE `C` and it should be updated (the rule is seen as certified in the future).

Routers drop expired packets, *i.e.,* packets containing an `exp_time` greater than the corresponding t_C. This ensures that in the absence of clock drift between the router and the RCE, and in the absence of router failure, a rule can be used for at most its lease period, *i.e.,* `exp_time`−`cer_time`.

To protect against clock drifts that cause t_C to increase faster than the RCE time, the router resets t_C (marks it as uninitialized) if it has not been updated for a predefined interval of time I (*e.g.,* several hours or days). In particular, the router resets t_C to the largest `cer_time` of any rule (certified by that RCE) that the router has seen during interval I. We assume that clock drifts between any router in the Internet and any RCE are bounded by a small value D within the interval I, such that D is much smaller than rule leases, *e.g.,* D is smaller than one second.

Finally, note that rules of unpopular RCEs could potentially be replayed by malicious hosts (unpopular RCEs are those whose certified rules do not travel for long periods of time through the network). However, we expect unpopular RCEs to not exist in practice, given that collocated destinations of one ISP should typically be certified by the same RCE, and thus all packets going to that ISP should use the same router timer. Moreover, even if unpopular RCEs would exist, given that no traffic is being sent to the particular destination

whose rules are replayed, the incentives for the attacker are small; more specifically, since no clients are accessing the destination, there are almost no incentives for DDoS attacks.

4.4 Anti-spoofing mechanism

If a source can spoof addresses on packets it sends, it can send packets to a destination `D` even if the rule does not allow it to, and in this way evade `D`'s policy. Moreover, one can mount a DDoS attack by using a single rule distributed by a malicious source to a set of colluders. To address this problem, RBF can use a previously proposed anti-spoofing mechanism. In this dissertation, we propose the use of ingress filtering, which is already deployed by over 75% of today's ASes [39]. When deploying RBF, RBF routers could also be used to apply ingress filtering. Note that if malicious ASes do not apply ingress filtering, DoS protection is not fully compromised as only hosts in these ASes can launch attacks.

Instead of ingress filtering, RBF could leverage other anti-spoofing mechanisms such as Passport [77]. However, Passport [77] requires a secure routing layer and incurs extra overhead in packets.

The anti-spoofing mechanism requires middleboxes and routers that change a packet's destination address also to change the packet's source address attribute.

Finally, note that RBF can function even with a partially deployed anti-spoofing mechanism, such as the currently existing ingress-filtering. The default-off nature of RBF scales down the bots available to attackers since it reduces the spreading potential of viruses (*e.g.,* viruses cannot perform random scanning because they have to know a rule for the contacted destination). Moreover, a DoS attack will stop once the lease of the rule expires; the victim can detect the attacker, and stop providing new rules to it for a period of time. To detect attackers in this case, destinations can use a rule to record routers on the path (*i.e.,* such as the the Pi [107] example in Section 3.3.) In addition, RBF could be used to incentivize the deployment of ingress filtering; destinations under attack can simply deny rules to requesters from ASes known not to ingress filter (*e.g.,* which can be learned through an independent service).

Chapter 5

Security Analysis

The RBF design aims to achieve the following three goals: (i) *policy enforcement* – ensure that the authorized rules respect the policies of all participants (routers, middleboxes, destinations), and packets with unauthorized rules are dropped inside the network; (ii) *rule enforcement* - rules cannot be used by malicious senders and, if senders or rule participants are untrusted, respect of rule directives can be enforced; and (iii) *rule safety* rules cannot be used to attack the network. Next, we summarize RBF's security properties, the threat model and assumptions under which they hold, and the mechanisms that allow RBF to meet these goals.

Assumptions: We assume that DNS resolution is secure, that distribution of RCE keys to routers is secure, and that RCEs are not malicious.

Attackers: An attacker in RBF can be any host, middlebox, or router: sources can attempt to attack destinations by forging, evading or tampering with their rules; destinations can try to attack the network by creating rules that waste resources and slow down routers; middleboxes and routers can attempt both of these attacks.

5.1 Security Properties

We decompose the aforementioned security goals into four specific desired properties:

1. **No Rule Forging:** A host `S` cannot manufacture a rule that sends packets to another host `D`, unless `D` explicitly agrees with this rule, *i.e.,* destinations and middleboxes control the creation of rules that send traffic to them.

2. **No Rule Tampering:** Sources, routers and middleboxes cannot tamper with the destination's rules.

\Mechanisms Properties\	Certifi- cation	Lease	Anti- Spoofing	Static Analysis
No Rule Forging	×	×		
No Rule Tampering	×			
No Rule Evasion		×	×	
Network Safety	×			×

Table 5.1: Properties and Defense Mechanisms

3. **No Rule Evasion:** Host `S` cannot send packets to destination `D`, if `D`'s rules do not accept packets from `S`.

4. **Network Safety:** A destination `D` cannot create unsafe rules. In particular, `D` cannot create rules that (a) cause infinite loops, (b) corrupt router state, (c) DoS routers or RCEs, or (d) violate ISP policies.

5.2 Mechanisms

RBF uses four mechanisms to achieve the above properties:

1. rule certification
2. rule leases
3. anti-spoofing
4. static analysis.

Next, we discuss how these mechanisms achieve the four aforementioned security properties. Table 5.1 summarizes this information at a very high level.

5.3 Analysis

Rule Forging

To forge a rule that sends packets to an address IP_D of a destination `D`, an attacker has to be able to receive packets addressed to IP_D as well as to send spoofed packets as originated at IP_D. The set of attackers in this category are mainly represented by the routers on the reverse path from the RCE to `D`. However, note that such attackers can already disrupt `D` at the routing level (*e.g.,* routers can drop/amplify packets to `D`, can insert malicious routing

updates such that packets do not reach D, *etc.*). Furthermore, using forging rules to mount a DoS attack is not easy. A malicious router creating a rule pointing to D would need to employ a set of colluders to which it can distribute the rule, and then have these colluders DoS the victim. In the absence of such colluders, the malicious router can only use the rule to send itself packets to D. Finally, the victim can detect the attack by checking whether it has authorized the rules used by the packets it receives. This also means that a malicious router M cannot deceit D into using rules created by M that allow M to eavesdrop packets sent to D, *i.e.,* rules in which M is imposing itself as an unwanted middlebox. If D does not recognize the rules it can immediately contact its ISP.

Based on these considerations, we believe that the possible rule forging attacks are relatively benign. Still, if needed, these attacks could be prevented by (a) having ISPs upload to the RCEs the binding between the IP and the public key of their customers, or (b) by having the ISPs themselves sign the binding between their clients' addresses and their keys (*e.g.,* through an extended DHCP like service). For example, in our presented approach, a large enterprise could protect against malicious routers forging rules to its addresses by having its ISP upload its public key to all RCEs, together with its address prefix.

Even though an attacker cannot forge a rule, it may still be able to send packets to a host which did not approve the rule. Assume the address of host D changes from $\texttt{IP}_D$ to $\texttt{IP}'_D$, and that $\texttt{IP}_D$ is immediately assigned to host X. If D grants a rule to host A, then A can use the rule to send packets to host X once the address has changed. However, we argue that the impact of such attack is negligible. First, the attacker can use the stale rule only for the duration of its lease. Second, the attacker cannot influence the selection of the new address owner, so it cannot pick its victim.

Rule tampering

The fact that rules are signed prevents any router or middlebox along the path to tamper with the rules.

Rule evasion

Host A can attempt to elude D's rules by (a) spoofing its address, (b) mounting a replay attack, or (c) using a reflection attack.

The anti-spoofing mechanism employed by RBF precludes an attacker (*e.g.,* A) from spoofing its address.

The lease mechanism mitigates the replay attacks, by limiting the ability of A to use the rule beyond the time intended by D.

Finally, an attacker A can indirectly send packets to a destination D by sending packets to a reflector R with a return rule pointing to D. D can defend against receiving reflected packets by specifying in each of its rules the sources that are allowed to send packets on that rule. Unless R is in this set, the routers will drop the packets reflected by R. A sophisticated

attacker might attempt to find rules that the victim D uses to communicate with R. If R is a large entity (such as Google), the attacker might have the opportunity to mount a DoS attack on D. However, to find such rules the attacker would typically need to collude with compromised routers on the paths between R and D, which can most likely already harm D as discussed earlier. In addition, R can detect whether the other communication endpoint does not send meaningful responses (the attacker does not receive any traffic back) and can detect such attacks.

Network safety

D can attempt to create rules which (a) cause infinite loops, (b) corrupt router state, (c) DoS routers, and (d) violate ISP policies. Furthermore, D may also attempt to (e) DoS the RCEs. Next, we consider each of these attacks in turn.

Infinite Loops – A rule could potentially create a forwarding loop by changing the packet destination in a circular fashion (the packets using such a rule would be dropped only when their TTL expires). To protect against such attacks, RCEs use *static analysis* to check whether the rule may cause forwarding loops. In Appendix A, we present an algorithm, called *Loop-Free Rule (LFR)*, that detects forwarding loops, and we prove the following result:

Theorem: A rule cannot create forwarding loops if it is validated by the *LFR* algorithm.

Similarly, in Appendix B we present an algorithm that detects rules that can lead to invocation loops, *i.e.,* the packet keeps invoking functionalities at one router.

Importantly, note that both these types of attacks (that form loops) can be avoided through runtime mechanisms (see Appendices A and B). However, we take advantage of the rule certification and keep the forwarding mechanism simple by using static analysis.

In a nutshell, to detect loops, it suffices to verify whether the finite state machine (FSM) modeling the rule execution has cycles containing one or more destination addresses. The states of the FSM are of the form <*packet.attributes*, *router.attributes*>. While given a rule it is easy to find the transitions between these states, the number of states can be large. Fortunately, the rule structure is very simple, and rules only allow comparison operations and assignments. This allows us to use a highly simplified version of symbolic-execution [75] to reduce the number of states to that of the number of destinations in the rule.

On the data plane, a malicious router can induce loops by changing the packet state. For example, a packet that has just visited a middlebox can be turned back to the middlebox by resetting its state to that before the middlebox. However, note that this attack does not represent an infinite loop since it requires the attacker to process the packet every time it comes back. For rules that filter out source addresses, the anti-spoofing mechanism can prevent most of these attacks because the router cannot send packets back to the middlebox while pretending to be the source.

Router corruption – The rule forwarding logic is simple enough so that we can assume that routers implementation of this logic is correct. Together with the fact that rules cannot

modify router attributes, this guarantees that rules cannot compromise router state.

Attackers might also try to corrupt routers by exploiting errors in the invoked router functionalities. However, routers can use techniques such as virtualization [12, 51] to protect against such attacks.

DoS Routers – With RBF, an attacker can attempt to slow down routers by sending packets with random certificates, causing the router to waste time verifying these bogus signatures. RBF routers can simply blacklist such attackers. Moreover, this attack can only occur at the first RBF router. For this reason, the incentives for it are low, as the attacker ends up targeting its own access route and, at most, other collocated users. The attack can further be alleviated by having two forwarding paths, a fast path for the already verified and cached certifications, and a slow path for first time rules. A more sophisticated version of this attack may use a large number of valid rules instead of random certificates. This attack is both more difficult to mount and to detect, and can be mitigated by placing a cap on the rate of new rules per host.

Another DoS attack on routers is to use rules that take a long time to forward. However, the rule forwarding time is directly proportional with the rule size, which is bounded.

Attackers may also attempt to slow down a router by extensively invoking the functionalities it exposes. In fact, there is a broader question about how one should architect router implementations to accommodate functions invoked by end users and, in particular, to make sure functions do not open additional vulnerabilities through resource exhaustion attacks. While a comprehensive treatment of this question is an active research topic [41, 53, 113] and beyond the scope of this dissertation, in Sections 6.1 and 6.2 we present and evaluate a prototype software router that offers isolation between the plain forwarding traffic and traffic invoking functionalities as well as isolation between different functionalities.

Path violation – Malicious hosts can create rules which violate ISP routing policies. For example, two colluders `A` and `B` could create a rule as follows: the packet leaves the source towards `A` but when reaching a certain AS or router (before reaching `A`), it suddenly changes the destination to `B`. In this way, with multiple colluders, arbitrary AS path could potentially be selected. To prevent such a malicious behavior, RCEs may use static analysis to guarantee that the destination attribute is not changed before the packet reaches the current rule destination.[1]

DoS RCEs – By sending certification requests at a high rate, attackers may attempt to indirectly attack other destinations by slowing down their RCEs. However, RCEs can control the number of destinations that use them and rate-limit the certification request for each destination (§4.1). Moreover, to improve availability, each host can use multiple RCEs. Malicious routers may also attempt to replay a destination's certification requests at a high-rate in an attempt to prevent it from certifying new rules. However, this does not represent a new attack ability for malicious routers, which can already drop any packet of

[1]Similarly to the case of infinite loops, this attack could be prevented by restricting the rule forwarding mechanism, but we again take advantage of the already existing certification.

that destination, just as they can also drop the certification requests.

Source Rules – Given that the control from the destination rule is never returned to the source rule, each packet containing a source rule can be considered as two distinct packets: one using the source rule and one using the destination rule. Thus, the use of source rules does not create any new threats compared to having a single rule in packets. Note that a source rule is just as any other rule and must have at least one destination (middlebox, indirection router, *etc.*). In order to prevent the path violation attack described before, the attribute that switches the control from the source rule to the destination rule can be set only when arriving at one of the destination addresses specified in the source rule.

Chapter 6

Evaluation

6.1 Implementation

This section describes our prototype RBF router and rule compiler.

6.1.1 An RBF Rule Compiler

Our prototype implementation offers users a high-level language in which to write rules, language that is largely identical to the rule syntax used in this dissertation. We wrote an RBF compiler in C++ that translates this high-level language into a compact rule format carried in packets. This compact format uses the following types of statements: `drop`, `sendto` address, `set` the value of a packet attribute, `compare` attributes (with values or between them), `jump` to a forward statement (always or on false/true conditions) and `invoke` a router-defined function. The `compare` statement sets a flag depending on the result of the comparison, which is used by the `jump` statements (the flag is initialized with zero at the start of the rule). The `jump` statements can only take positive offsets. To reduce the size of rules, we use variable-length encoding in representing the internal rule structure. We also use: 8B(ytes) for public-key hashes, 3B for the user-local index (contained in the rule identifier), 3B to identify the RCE, 3B to identify router-defined functions that do not require approval to be invoked and 8B for those that do, and 2B as the default RBF packet attribute values.[1] For the lease we use an absolute expiration time consisting of first 4B of the NTP format, with second-level granularity and a wrap-around period of 136 years. The maximum rule description size is 256B in our implementation.

[1] Our current prototype only supports this default size.

6.1.2 A Prototype RBF Router

Rationale: We implemented RBF forwarding using Click [83] and RouteBricks [48]. Most commercial routers implement packet processing using ASICs or specialized network processors (NPs) rather than general-purpose CPUs and, as such, our software-based prototype is not entirely representative of currently deployed routers. To a large extent, our choice of prototyping platform is borne of necessity since commercial routers are closed. Beyond necessity, however, we believe a software-based prototype is valuable for multiple reasons. First, recent research [48,50,65] has demonstrated that, with modern multi-core servers, it is now possible to build high-speed software routers up to edge and even core speeds. Secondly, while not directly reusable, several aspects of our implementation architecture such as our approach to partitioning tasks across multiple cores should apply to network processor-based routers. Finally, several research [23, 60] and commercial switches [8] augment ASIC-based switches with some number of co-located general-purpose cores or servers for greater flexibility in packet processing – our prototype architecture is directly applicable to such platforms.

Design requirements: We build our prototype in the context of modern multi-core servers that incorporate multiple processors or "sockets", each with multiple cores [5, 29]. As shown in Fig. 2.1, the software stack of an RBF router includes the following key components: (1) an IP forwarding module, (2) the rule execution engine, and (3) some (possibly zero) number of specialized forwarding function modules. All packets traverse the rule execution and IP forwarding components, while different subsets of packets may traverse one or more specialized functions. In addition, the resources required to process a packet may vary widely across functions; *e.g.,* an encryption function would use lots of CPU but little cache, while a caching module may use more cache and less CPU. At a high level, our design goal is to balance high performance (*i.e.,* making efficient use of resources) with performance isolation, both across different functions, and between functions and the rule execution engine (*i.e.,* sharing resources in a fair manner).

Approach: In its full generality, the above goal requires contention-aware scheduling that simultaneously takes into account the multiple resources (cores, various caches, memory bandwidth, I/O bandwidth) for which tasks might contend. For modern multi-core systems, this is in itself an area of active research [41, 113] and beyond the scope of this dissertation. Instead, in our prototype, we address the issue as follows. The IP forwarding module and the rule execution engine are the central, most critical, components of the router and hence we assign these to a socket of their own and do not run specialized functions at cores in this socket. This avoids having the IP and rule execution engines contend with specialized functions for cache, CPU and other resources at the cost of some potential inefficiency since these "reserved" cores (if unused) cannot be used by specialized functions (if needed). We then assign specialized functions to the remaining "unreserved" cores. We rely on the existing (Click and Linux in our implementation) system schedulers to ensure fair sharing of CPU

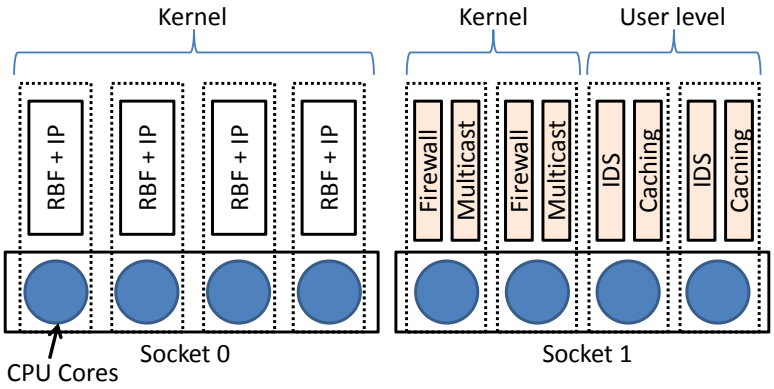


Figure 6.1: Core Allocation Example in RBF Router

resources between functions on the same core.

To achieve high performance, we run a single thread performing both IP forwarding and rule execution at each of the reserved cores; this ensures that packets that do not invoke any specialized functions are processed entirely by a single core avoiding potentially expensive cache misses and inter-core synchronization [48]. Packets that invoke specialized functions must be relayed across cores and hence incur corresponding performance overheads due to cache misses and so forth. To improve the efficiency of such transfers when these functions are implemented in user space, we use shared memory pages and event queues. An example of the resulting system architecture is depicted in Fig. 6.1.

6.2 Evaluation Results

We use our prototype to evaluate the overhead RBF imposes on packets (§6.2.1), routers (§6.2.2) and RCEs (§6.2.4).

6.2.1 Packet Size Overhead

Fig. 6.2 presents rule sizes (in bytes) for a range of examples, including those from §3.3. The figure captures all the RBF-related fields and presents the size broken down into (a) the rule and the associated attributes' binary encoding; (b) the control fields used for the lease, RCE identification, to specify whether the return rule is in the packet, whether the packet has a source rule, and so forth; and (c) the rule signature. We assume a 41B signature obtained using ECDSA with ECC public keys for RCEs derived from the NIST B-163 or K-163 curves [30], offering 80 bits of security. Note that RBF is independent of the exact signature scheme used and that smaller (and faster) signatures can be used. However, shorter RCE keys may require more frequent updates to compensate for the lower security

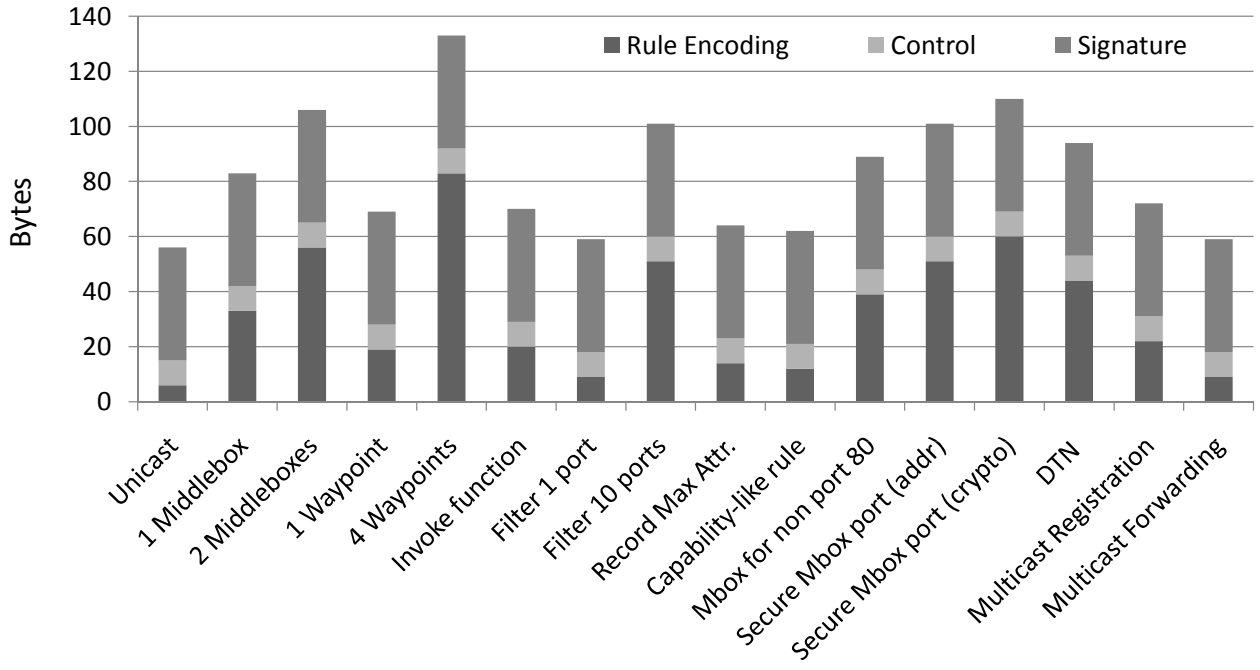


Figure 6.2: Rule Sizes

guarantees. The rules in Fig. 6.2 do not contain an identifier, and are identified by endpoints and routers using a hash over their content. Rule identifiers are required for rules whose content may change during a communication (such as the rules of mobile hosts) and incurs an additional 11B overhead in our implementation (8B for the hash of the public key and 3B for the user-selected index). Note that the rule identifier need be unique only with respect to a single communication endpoint (*i.e.,* all parties that a host `X` communicates with should have unique rule identifiers).

From Fig. 6.2 we can see that many common forwarding scenarios (unicast, routing via middleboxes, rules for DoS protection) can be expressed with around 60-80B rules while more complex rules (*e.g.,* loose source routing, secure middleboxes, anycast) can take as much as 140B. The average rule size across all examples we have implemented is 85B, representing 13% overhead for an average packet of 630B [9] and 6% overhead for a 1500B packet. By comparison, using RSA-1024 signatures (instead of ECDSA) would incur 27% overhead on a 630B packet and 11% overhead on a 1500B packet. Note that the packet size overhead matters mainly when links are congested, in which case the goodput (application-level throughput) of the participants to the traffic is reduced proportionally with the RBF overhead.[2] Links usually get congested when many endpoints use throughput-intensive applications. Since throughput-intensive applications typically use large packets, we expect

[2]The added latency due to the additional per-packet overhead is minimal.

the observed goodput reduction to be in practice closer to the overhead associated with large packets rather than average-sized packets. Moreover, we expect this overhead to be significantly reduced in the future with the adoption of jumbo frames.

Potential Optimization - Rule Caching: Per-packet overhead can be significantly reduced by *caching* rules at endpoints and routers; packets whose rules have been cached need only carry rule identifiers. There are two opportunities for caching. First, destinations can cache return rules; this allows the return rule to be eliminated from all but the first packet in a source-to-destination exchange. Second, rules can also be cached at routers. Here, however, we must ensure no packet carrying only a rule identifier arrives at a router that does not store the corresponding rule description. This might occur, for example, due to a route change or when a router deletes the rule from its cache. In such cases, the router can simply drop the packet in question, if the endpoints include the rule on all retransmissions and during periods of high packet loss. Of course, caching imposes additional storage overhead at routers as we evaluate shortly.

In summary, based on our evaluation, we see that the per-packet overhead due to RBF can range from as low as 24B when using caching and up to ∼250B in the bad case where there is no caching and the packet carries complex destination and return rules. Sources can add their own control for outgoing packets through source rules, at the expense of an additional rule in the packet. In general, the more endpoints take advantage of the flexibility offered by rules, the higher the overhead. Based on their interests, endpoints can select different points on this tradeoff between flexibility and overhead.

6.2.2 Router Overhead

In this section, we evaluate the overhead RBF imposes on routers for rules that do not invoke specialized processing functions; we consider router functions in the following section. The primary overhead RBF imposes on routers is the additional processing required to execute and authenticate rules and the additional storage capacity required if rules are cached.

Rule Forwarding

We first measure the overhead of rule processing by comparing the performance of RBF-on-RouteBricks to that of unmodified RouteBricks running on a single high-end server machine. We use a dual-socket server with four 2.8GHz Intel Xeon (X5560) cores per socket to (from) which we generate (sink) traffic over two dual-port 10G NICs. In this experiment, we use all 8 cores to forward packets and do not perform rule authentication (*i.e.,* the RBF router is assumed to be inside a trusted domain).

Fig. 6.3 plots forwarding rates for some of the examples from Fig. 6.2. The first column represents a packet stream with sizes generated based on a packet trace collected on the Abilene backbone [22]; since the packets from the trace do not have rules, we add to each

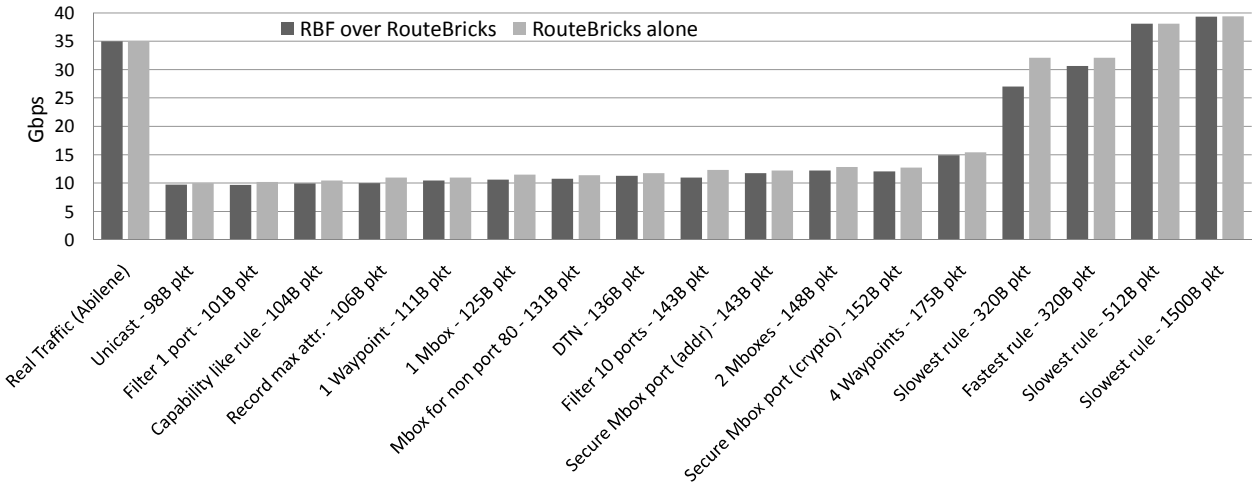


Figure 6.3: Forwarding speed for RBF over RouteBricks

packet the slowest rule that fits in the packet. By "slowest" we mean the rule that takes the longest time to forward, as determined by the number of conditions and actions encountered during forwarding. To capture the performance impact for small packets, we profile each rule without any payload and with no return rules. In the figure, packet sizes are shown next to the example name and entries are sorted in order of increasing packet size; the packet size also includes the Ethernet and IP headers. The last columns depict forwarding of larger packets, *i.e.,* that also contain data payload. To see the impact of the type of rule for these packets, we profiled them with the fastest and the slowest rules. Note that all rules are profiled in the worst case, meaning that the longest path through the rule is considered. For the slowest rule we use a 145B anycast rule which selects one out of 10 destinations based on the value of a packet attribute (the last destination is always selected in our tests to use the longest path in the rule).

Overall, we see in Fig. 6.3 that the performance degradation due to RBF's more complex per-packet processing is always modest (<15%) and virtually non-existent at larger packet sizes. For small packets the CPU is the forwarding bottleneck, and RBF's added processing slows the router. For larger packets the I/O system is the bottleneck, and there are enough free CPU cycles to execute rules. A fine-grained profile of the rule execution module showed that it uses between 120 CPU cycles per packet for the fastest rule and 600 CPU cycles for the slowest rule; in comparison, the IP router used in our experiments requires around 3000 cycles per packet without rule execution. Also note that compared to the network-level forwarding results from Fig. 6.3, application-level goodput is further reduced by the RBF header.

RBF is also compatible with high-end routers that use specialized hardware, rather than the general-purpose x86 hardware we consider. Increasingly, high-end routers rely on network

processors [49, 103] for data plane processing (*e.g.,* Cisco's flagship CRS router uses network processors). We expect network processors can accommodate the flexibility RBF requires in per-packet processing (since network processors are fully programmable, much like our general-purpose server but simply use a different programming paradigm) with a similar type of overhead.

Rule Authentication

Recall that routers must verify rule signatures. While signature verification is expensive, there are two reasons we argue this overhead is manageable. First, only routers at trust boundaries need to authenticate rules; border routers see lower traffic loads than core ones, and hence can more easily accommodate the cost of authentication. Second, routers can cache the results of authentication checks, maintaining a hash of the rule and signature. With caching, the full signature verification is only required for the first packet forwarded on a new rule, within a certain caching period. Thus verifications can be limited to only border routers and, assuming a large enough cache, the rate of verifications is on the order of the rate of arrival of packets with new rules.

In this dissertation we proposed the use of ECC signatures applied on rules and we are pursuing a hardware based implementation of RBF that uses ECC signatures. In the meantime, we have evaluated RBF with a software implementation of RSA which is more amenable to software implementations (we have compared the packet overheads of ECC and RSA previously, in Section 6.2.1). New CPU hardware can handle thousands of public key signature verifications per second. For example one core of the router machine that we used to forward rules in the previous experiment can perform over 38k verifications per second for RSA 1024. Moreover, these verifications can be parallelized; if all the 8 cores of the router machine are used, the machine can process almost 300k 1024bit RSA signatures per second. However, if the same hardware is used both to verify signatures and to forward packets, the extra processing required by authentication may slow down forwarding.

In Fig. 6.4 we evaluate the impact of rule verification on forwarding, using the same two machine setup as before. On the X axis we show the percent of packets with rules being authenticated. For the rest of the packets (which are not authenticated), we assume the router has cached authentication information, since those packets belong to flows already established. Note that we ignore the cache lookup time. We use 1024 bit RSA keys. To decouple the effect of rule processing from that of authentication, we profiled different packet sizes containing the slowest rule; to capture the impact of authentication for small packet, we also profiled this rule with no data payload. The vertical line at approximately 5% in Fig. 6.4 shows the average rate of packets corresponding to new flows in the Internet; we extracted this rate from recent backbone traffic monitoring at [9][3] However, we think this

[3]This ratio represents the number of flows to the number of packets (*i.e.,* the inverse of the flow length in packets). We use this value in this experiment because we fill up the entire available bandwidth. As mentioned in §3, we have observed a rate of less than 1% when comparing the bandwidth used by packets

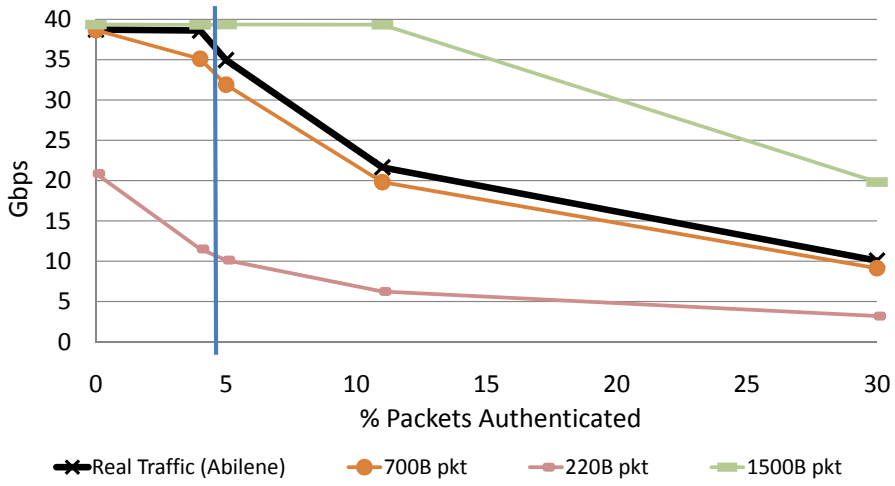


Figure 6.4: Forwarding speed when verifying packets.

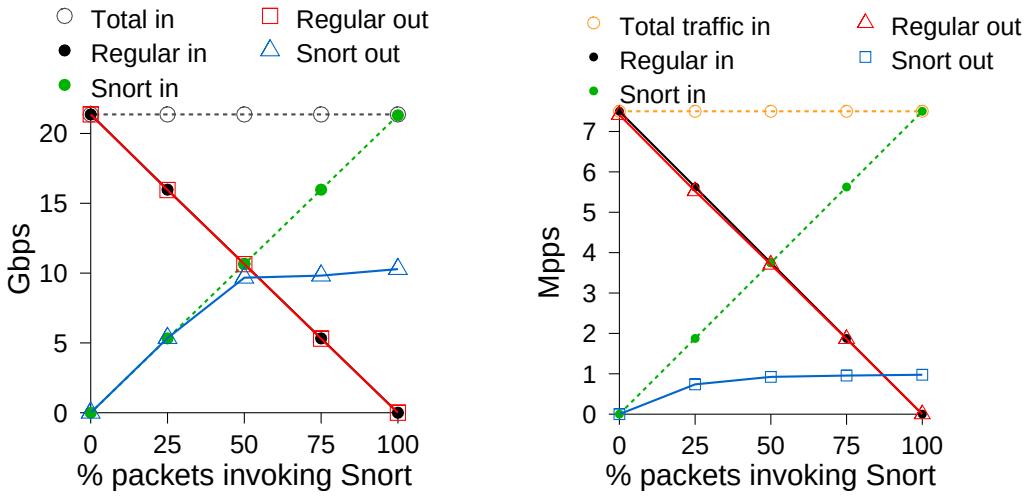
rate represents a pessimistic estimate for average rate of packets with new rules seen by a router because multiple connections share the same rule, *e.g.,* a client makes multiple connections to a web server all sharing the same rule.

As we can see, authentication does impact forwarding as the CPU becomes the scarce resource. As expected, this impacts mostly the smaller packets. However, the rate for the packets generated using real traffic data (Abilene) is not significantly impacted up to the 5% threshold derived from the Internet's average rate of new flows. Moreover, the good news is that by adding more cores, this process can be parallelized; current industry trends seem to indicate that end-hosts as well as routers will increase their number of cores. We thus expect a linear increase in performance as more and more cores are used, up to the forwarding results presented earlier in Fig. 6.3.[4]

Slowing down routers by purposely sending many valid new rules is difficult. First, the attacker has to find a large number of valid rules and use them before they expire. For example, assuming a 1 million rule cache, the attacker would have to find over 1 million valid rules. Second, routers can detect such behavior and block attackers, similar to the attack with random certifications described in §5. For example, a typical user is expected to send packets containing only a few new rules per second. In addition, as mentioned in §5, the incentives to mount such attacks are not as high as, for example, DoS attacks, since attackers would have a short range of action. Essentially, attackers can only attempt to slow down their access path and attack their own provider.

indicating new flows to the link capacity, because the link is not fully utilized.

[4]In addition, historically, network speed has increased at a smaller pace than computational power (see Nielsen's law [18] vs. Moore's law); accordingly, we expect such a trend to persist in the future, in which case, authentication will be more and more easy to perform at line speed.



(a) Large Packet Trace

(b) Small Packet Trace

Figure 6.5: Performance Isolation Between Router Functions and Regular Traffic

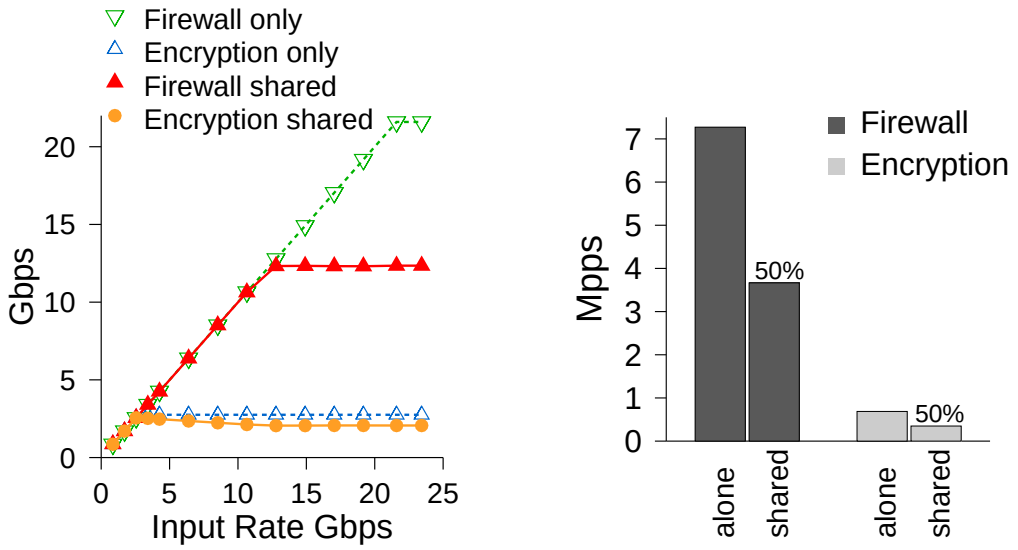
Router cache sizes

We earlier proposed that routers cache rule authentications and/or even rule descriptions. In each case, the number of cache entries required depends on the number of distinct rules the router sees. If we assume that all packets in a flow share the same rule, then the number of distinct rules passing through a given router varies between the worst case of O(#flows) to the best case of O(#destinations) seen by the router. The former corresponds to a destination that uses a different rule for every source it communicates with, the latter to a destination that uses a single rule for all potential sources.

In our implementation, each cached authentication is 19 bytes – 11B for the rule identifier and an 8-byte hash value used to verify whether the rule has changed since it was authenticated. Each router uses its own secret hash function to prevent attackers from using hash collisions. Thus, one million rules would require only 19MB of memory. For caching entire rules, Fig. 6.2 reveals average and worst-case rule sizes of 85 and 133 bytes, respectively. If we conservatively assume traffic is uniformly distributed across these forwarding categories, we arrive at an estimated cache size of 85MB (average) to 133MB (worst-case) for 1M rules, which is within the scope of memory available in current routers.

6.2.3 Router Functions

Our router prototype supports specialized functions implemented at either kernel-level or user-level. We currently support three router functions: (i) the Snort IDS [24] adapted to



(a) Large Packet Trace

(b) Small Packet Trace

Figure 6.6: Fairness Between Router Functions

run as a user-level function, (ii) a kernel-mode firewall implemented in Click and (iii) a kernel-mode encryption engine also implemented in Click. Each function runs as a separate process/kernel thread isolated from the packet forwarding path through queues. We measure performance and fairness using the above functions on the same hardware as before. We dedicate four cores to the standard forwarding path and the remaining four cores to custom functions.

Fig. 6.5 illustrates the resource isolation between forwarding and router functions; the function used in this experiment is Snort (running on four cores). To generate traffic we use real traces of malicious traffic created particularly for IDS testing [20, 64]. The average packet size of the trace used in Fig. 6.5(a) was 1065 bytes while in Fig. 6.5(b) was 195 bytes. To avoid biasing our results, we modify Snort not to drop any malicious packets so packets are only dropped due to resource exhaustion. Our test maintains constant total input traffic while increasing the percentage of input traffic that invokes Snort (X-axis). We see from Fig. 6.5 that Snort traffic does not affect the "regular" traffic that does not invoke Snort, in the sense that no regular traffic is dropped, even as a growing percentage of input Snort traffic is dropped.[5]

Fig. 6.6 illustrates isolation between router functions. We use two functions, firewall

[5]We also measured the performance of the system with all the eight cores running both forwarding and Snort, and all the packets directed to Snort. While this configuration does not provide isolation for the regular traffic, it can forward a higher total throughput of 22Gbps when using large packets.

and encryption, and we run three experiments: (1) all traffic invokes the firewall function and no traffic invokes encryption; (2) all traffic invokes encryption; and (3) equal halves of traffic invoking firewall and encryption. In the third (shared) test the CPU is shared fairly between functions (we use Click-level scheduling); thus, the ratio between the maximum throughputs achieved by each router function is expected to roughly match the ratio between the throughputs of the functions when running in isolation. Fig. 6.6(a) plots the resulting forwarding rates under increasing input traffic, when generating packets from a trace with an average packet size of 1065 bytes. Perhaps surprisingly, in Fig. 6.6(a) the encryption throughput is higher for a mix of firewalled and encrypted traffic than 50% of the throughput when encryption is executed alone. The reason is that the trace contains large packets. In this case, the CPU is not the bottleneck for the firewall functionality, but is the bottleneck for encryption (since it is more CPU-intensive), and thus encryption ends up using the leftover firewall CPU cycles. Fig. 6.6(b) shows this result when using a trace with very small packets – the average size is only 195 Bytes. In this case, both functionalities achieve around 50% of their throughput in isolation. In general, the high throughput achieved when running each function in isolation illustrates the benefit of running instances of a single function at multiple cores, as opposed to one function per core, since this allows the unused resources from one function to be seamlessly utilized by other functions.

6.2.4 RCE Load

We use a simple back-of-the-envelope calculation to estimate the total number of RCE servers required for the Internet. The bulk of requests to RCEs are determined by IP address changes and per-client certifications requested by sites that protect against DoS (by redirecting DNS requests to powerful entities, see §3.3,§4.2). Note that in the latter case, requests to RCEs are made only for approved customers. There are currently around 800 million hosts in the Internet [31]; for simplicity, we consider 1 billion hosts. We assume a worst-case scenario in which all hosts request certifications in the same second, *i.e.,* these requests are made either by hosts individually or by DoS-concerned websites that hosts are trying to access. We implemented RCE rule certification in software using RSA signatures, and measured it on the same 8-core server used throughout our evaluation. We find a single server can achieve a certification rate of over 16,000 rules per second. Based on benchmarks of our implementation and assuming an oversubscription rate of 10× (ISPs today commonly oversubscribe by 100×), the total load due to certifying rules above could be accommodated by around 6,000 servers; *e.g.,* handled by 20 RCEs with 300 servers each. Hardware implementations might reduce this number by more than an order of magnitude. For example, using recent ECC prototypes [70, 112] a single ASIC could potentially perform 40,000 RCE certifications per second, requiring a total of only 2500 such devices.

Chapter 7

Related Work

RBF is inspired by and extends several directions in past research. At a high level, RBF's contribution is in offering extensive flexibility while respecting policies, where prior approaches tended to focus on one or the other. We next contrast RBF to a large body of work proposing more flexible or more secure network architectures. At the end of this section we present a table summarizing most of the properties by which RBF differs from other proposals. We then also show that key to RBF's ability to provide both flexibility and policy compliance is a new division between the data-plane and control plane compared to previous work.

RBF's focus on flexibility in forwarding is, of course, similar to that of active networks [101]. Where RBF differs is in constraining this flexibility—RBF forwarding directives are simple `if-then-else` constructs rather than arbitrary code. And while RBF does accommodate specialized functions at routers, we assume these are installed by network providers. Nor does RBF permit endpoints to introduce or manipulate router state (unlike, for example, schemes such as ESP [42]). Because of this constraint, RBF cannot support as complex forms of packet processing – *e.g.,* RBF cannot support the forms of reliable multicast and multicast feedback thinning that ESP does. The flip side is that, because of RBF's constraints on rules and because RBF mandates the use of these rules, it achieves greater security through provable rule safety and the ability to prevent unwanted traffic by default.

More generally, previous architectures that aim to provide flexible forwarding *e.g.,* [37, 42, 78, 86, 90, 98, 101, 104, 106] can be divided into four classes based on whether end-users are allowed to (i) modify router forwarding state, or/and (ii) modify forwarding information in packet headers (*e.g.,* destination address):

1. Modify both router forwarding state and forwarding information in packet headers (*e.g.,* most Active Network proposals [101], ESP [42]).

2. Modify router state but not packet headers (*e.g.,* Active Networks focusing on "active storage" [102]).

3. Modify packet forwarding information but not routing state (*e.g.,* i3 [98], DOA [104]).

4. Modify neither router forwarding state nor packet state (*e.g.,* IP). [1]

RBF belongs to the third class above, which we argue provides the best tradeoff between flexibility and security. The first two classes allow data packets to modify router forwarding state, which poses significant security risks. Indeed, at the limit, an application could implement complicated distributed protocols (*e.g.,* routing protocols) whose safety is notoriously hard to verify. In contrast, the last class offers limited flexibility, as end-users can exercise no control on packet forwarding.

RBF is likewise inspired by the growing body of research advocating a default-off model for network connectivity [38, 43, 66, 72, 91]. What RBF adds is a focus on flexibility (unlike, for example, [38, 66] that only consider simple access control), extending this flexibility to the wide-area (unlike [43, 72] that focus on enterprise or data-center environments) and in providing forwarding directives with end-to-end semantics (unlike [91]). Moreover RBF achieves the above without requiring the dynamic creation/teardown of state at routers (unlike [38]) or centralized policy entities (unlike [43, 72]).

RBF also shares aspects of the work on capability-based architectures (*e.g.,* [87,108,110]) since rules are essentially capabilities. RBF uses capabilities/rules in all packets and (again) adds a focus on flexibility in forwarding, showing how endpoints can choose to incorporate middleboxes and router extensions, implement multicast, anycast, multipath routing and so forth. RBF also differs in the mechanisms adopted. In particular, RBF does not require different router-level treatment for control traffic, opting instead for a scalable distributed infrastructure to alleviate DoS attacks on the control plane.

More generally, there are a vast number of proposals for DoS protection [37,47,68,78,79]. RBF offers similar protection while introducing a focus on flexible forwarding. Fundamentally, RBF's ability to offer flexibility is because rules are not bound to in-network state (unlike, for example, [37] or [68]).

RBF also draws inspiration from overlay-based architectures that aim to extend IP forwarding [34, 71, 98, 104] but instead operates at the network layer. This offers improved security (since RBF can block unwanted packets even if the attacker knows the target's IP address) and the ability to influence path selection at the router-level.

NUTSS [61] employs on path middleboxes to provide control access, and off-path middleboxes to implement per-domain on/off policies. Compared to NUTSS, rules in RBF are global, which allows RBF to block traffic closer to the source. In addition, RBF provides a more flexible data plane, and does not require inter-domain synchronization to support functionality such as mobility.

RBF also adopts the philosophy of decoupling forwarding and routing proposed by path-pinning architectures such as Platypus [90] and SNAPP [86] but extends these for greater

[1] Of course, routers perform logging and monitoring tasks which change their state, and routers modify the TTL value which represents state in packet, but here we refer to *user-controlled forwarding* abilities.

	Flexibility Available to End-hosts						Security / Policy Compliance				
Property / Architecture	Explicit Middle-box Support	Multi-ple Paths	Invoke Router Extensions (*e.g.* IDS, multicast)	Use Router State in Forwarding (*e.g.* anycast, DTN)	Record Router State (*e.g.* network probing, ECN)	Mobi-lity	Policy Compliant Loose Paths	Policy Compliant In-network Functionality Use	Receiver Reacha-bility Control	Host DDoS Protec-tion	Safety of Network & Routers (*e.g.* loops, break ISP policies)
RBF	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**	**Yes**
Active Networks	Yes	Yes	Yes	Yes	Yes	Yes	No	No	No	No	No
ESP	No	No	Yes	No	Yes	No	No	No	No	No	Yes
i3, DOA	Yes	No	No	No	No	Yes	No	No	No	No	Yes
Platypus, SNAPP	Yes	Yes	No	No	No	No	Yes	No	No	No	Yes
TVA, SIFF	No	No	No	No	No	No	No	No	No	Yes	Yes
NUTSS	Yes	No	No	No	No	No	Yes	No	Yes	No	Yes
PushBack, AITF, StopIt	No	No	No	No	No	No	No	No	No	Yes	Yes
Predicate routing, Off-by-default	No	No	No	No	No	No	No	No	Yes	No	Yes
ICING	Yes	No	No	No	No	No	Yes	No	Yes	No	Yes

Figure 7.1: A comparison of RBF to: Active Networks [101,105], ESP [42], i3 [98], DOA [104], Platypus [90], SNAPP [86], TVA [110], SIFF [108], NUTSS [61], PushBack [68], AITF [37], StopIt [78], Predicate Routing [91], Off-by-default [38], ICING [93]

generality in forwarding.

Karsten *et al.* [74] propose an axiomatic model for network communication. The forwarding primitives used in [74] differ from rules in two ways. First, RBF can modify the packet header (attributes) in a more general way than the axiomatic model that only supports packet encapsulation. Second, RBF does not allow packet replication for safety and manageability reasons. Note that both RBF and [74] do not enable the forwarding process to modify state at routers.

RBF is complementary to recent efforts proposing open router APIs [27, 28, 81] – we offer an overall network design by which endpoints use the new functionality these router architectures promise to enable.

Comparison Table: Fig. 7.1 presents a summary of comparing various flexibility and security properties of RBF with those of several previous proposals: Active Networks [101, 105], ESP [42], i3 [98], DOA [104], Platypus [90], SNAPP [86], TVA [110], SIFF [108], NUTSS [61], PushBack [68], AITF [37], StopIt [78], Predicate Routing [91], Off-by-default [38], ICING [93]. Each column in Fig. 7.1 represents a property related to flexibility and/or security. For example, only the traditional active network proposals, ESP and RBF are architectures that expose router extensions (*i.e.,* router-defined functions) to end users, allowing them to invoke these extensions; however, only RBF guarantees that the use of these in-network

functionalities is policy compliant. Note that by reachability control we refer to destination's ability to specify which packets can reach it and which packets should be dropped in the network.

Data-plane vs. Control-plane division: A key feature that distinguishes RBF from previous proposals and allows it to achieve both flexibility and policy compliance is its division of functionality between the data and control planes. Active Networks typically make little use of the control plane, as they deploy the forwarding functionality and enforce security on the data plane. This makes policy compliance hard to achieve. In contrast, more recent proposals such as OpenFlow [23] rely heavily on the control plane and install flow state in the network to make sure the data plane respects the appropriate policies. This approach, while simplifies the data plane, results in a more rigid architecture. For example, supporting host mobility and traffic engineering require tearing down the old paths and instantiating new ones. These are expensive operations which have a negative impact on the scalability of these proposals. In contrast, with RBF, each packet contains (in its rule) enough information to prove to routers that it respects the policies of all participants involved in forwarding the packet. RBF achieves this property despite the fact that neither the routers nor the packet contain the policies. Thus, RBF retains the datagram model of the IP, unlike other recent proposals (*e.g.,* network capabilities [108, 110], ICING [93] and OpenFlow [23]), which are more akin to a connection-oriented model.

Finally, in RBF we mainly focus on the flexibility and policies available to end-points rather than to ISPs and, as a consequence, rules are not tightly coupled with routing algorithms, *e.g.,* unlike in ICING [93], NIRA [109] or Pathlets [57]. We have taken this decision because forwarding flexibility is typically associated with end-points rather than ISPs, and because ISPs are likely unwilling to directly involve end-points in the routing decisions. ISPs can reuse RBF to achieve their own desired flexible forwarding patterns by defining themselves the rules of end-points (rules could be returned by DHCP) or by "rule tunneling" (*i.e.,* stamping the ISP rule on the packets).

Chapter 8

Discussion and Limitations

RBF is an architecture that we argue strikes a desirable balance between flexibility and the ability to guarantee policy compliance of all network entities. We started this work with two high level goals in mind. First, we wanted a *complete* architecture that supports not only previously proposed communication primitives, but also future ones. Second, we wanted an *efficient* architecture in that a packet unwanted by a receiver along its path is dropped as early as possible.

While completeness in this context is difficult to formalize, intuitively we have reduced it to (1) supporting arbitrary communication paths, and (2) allowing all network entities (*i.e.,* sender, receivers, middleboxes, and routers) to be involved in the decision process. In other words, we wanted to be able to define virtually any forwarding path and give all involved parties a say in defining it. We noted that such a path can be encoded by associating with each node an "if-then-else" code snippet, which specifies the next node down the path. We further noted that allowing different network entities to define the communication pattern is equivalent to allowing them to define these code snippets. This is roughly what the RBF proposal is.

These goals are ambitious – they subsume, unite and extend many years of proposals for greater flexibility and security in networks – and much of RBF's complexity follows from these goals.

In addition to RBF's complexity, there are a number of aspects concerning RBF's design that might raise concerns such as: Can RBF be used to violate the polices of some of the ASes? Can RBF be deployed incrementally? What are RBF's limitations? In the rest of this section we try to answer these questions by discussing: RBF's definition of policy compliance and its implications (Section 8.1), incremental deployment paths for RBF (Section 8.2) and RBF's limitations (Section 8.3).

8.1 Discussion on Policy Compliance

A rule is explicitly guaranteed to comply with the policies of all the entities named in the rule (in fact the policies of the organizations owning these entities). The entities named in the rule are: (1) the waypoint routers, middleboxes and invoked functions, which enable the rule's flexible forwarding and (2) the destination. This property enables RBF to address the security concerns associated with previous proposals for more flexible forwarding such as source routes and active networks, making sure that any extra flexibility can be fully constrained by the policies of the participants; this property also enables destinations to protect against DoS attacks.

However, rules do not name all the entities participating in forwarding a packet. Thus, a legitimate concern is whether rules respect the policies of all the ASes on the path. More specifically, there are two distinct concerns that we discuss next.

First, some regard the current IP forwarding layer as not policy compliant, since it does not guarantee that the agreed path (given by routing) has been respected by ASes [93]. In this context, one might question whether RBF is able to enforce stronger guarantees for policy compliance than the current IP. In short, the answer is no. RBF operates on top of a routing-controlled point-to-point forwarding layer and is not concerned with how this forwarding occurs. In a different design, RBF's mechanism could (in theory) be applied to enforce path-compliant forwarding by specifying a waypoint in each AS along the desired path (and potentially using cryptographic functions to prove that the path was obeyed). However, we leave such exploration to future work. We note that RBF can run on top of any improved (policy-compliant) version of a point-to-point forwarding layer.

Second, due to its ability to use loosely specified paths (through waypoints), there is the question of whether RBF can violate the policies of some ASes. For example, ASes between two waypoints may not want to relay traffic from the original source to the final destination. We argue that this is equivalent to the question of whether overlay networks violate ISP policies or not, *i.e.,* waypoint routers/middleboxes can be seen as participating in an overlay. Although RBF belongs to the network layer, rules cannot change the state at routers and, compared to overlays, the only mechanism added by rules that can affect ASes is the ability to invoke router functions. However, function invocation is fully policy compliant, since each AS can restrict which rules can invoke its functions. The debate of whether overlays violate ISP policies is arguably an open question, which still generates powerful debates in the community and we leave a comprehensive discussion on this topic to future work. Our position in this debate is that as long as overlay participants agree on a loosely specified path, the lower level point-to-point forwarding cannot and should not distinguish this traffic from other traffic occurring between the waypoints, because the waypoints pay for the relayed traffic as if that traffic was received/generated by them.

Finally, to offer stronger policy-compliance guarantees, one could combine the RBF layer and the routing-controlled forwarding layer into a single layer. This is a valid point that can be applied to any monolithic vs. layered design, and we leave its exploration to future

work. The principles and mechanisms of RBF can be used to guide the design of such an approach. We note, however, that such a layer would be more complex, in that (1) the control plane is concerned with both routing as well as the enhanced flexibility offered through in-network functions, middleboxes, multiple paths, mobile hosts, *etc.* (for example a distributed implementation of such an algorithm in a similar way to BGP would be quite difficult) and (2) the data plane would be more heavyweight, as it has to always contain information about both the policy-compliant flexibility as well as the policy-compliant routing.

8.2 Incremental Deployment Through Infrastructure Upgrade

RBF requires upgrading routers, DNS servers and end-hosts, but is nonetheless amenable to incremental deployment through strategies similar to those proposed by prior clean-slate proposals [46, 55, 76]. In this section we discuss strategies for deploying support for RBF in the network infrastructure.

Initially, only a part of the infrastructure and a part of the traffic could be RBF-enabled. Indeed, all the benefits of RBF shown in Fig. 7.1 except receiver reachability control and DDoS protection can be achieved with a partial deployment of RBF routers and middleboxes. In a partial deployment, not all routers within an AS are upgraded to handle RBF packets. For example, an AS might start by upgrading only border routers and a selected part of the other routers and middleboxes. Such an arrangement should suffice for a wide category of uses of rules (*e.g.,* to leverage middleboxes, for filtering, path selection, mobility, *etc.*) although it is non-ideal for rules that require the participation of every router along a path (*e.g.,* a rule that records the worst-case congestion along a path).

In an initial deployment phase, RBF routers could support both RBF and legacy (non-RBF) traffic. Techniques such as network slicing and virtualization (*e.g.,* [19, 45, 94]) can be used to enable RBF traffic to coexist with legacy one. We also note that rules can always incorporate RBF-aware middleboxes located in ASes that do not support RBF. To also offer DoS protection and reachability control, individual ASes can fully upgrade to RBF by dropping legacy traffic.

Until all ASes are upgraded to RBF, some hosts in ASes that have already upgraded to RBF may want to communicate with hosts in ASes using legacy traffic. A simple way to achieve this is through multihoming, *i.e.,* have two interfaces one to an RBF-enabled AS and one to a legacy AS. However, multihomed hosts will be vulnerable to DoS attacks on legacy interfaces. An alternative to multihoming is to have RBF-enabled ASes support the communication with legacy ASes by encapsulating/decapsulating packets at their borders. To allow RBF-enabled and legacy end-hosts to communicate, an RBF-enabled AS would know its customers' rules and have its border routers actively add rules on packets arriving from a legacy host destined to one of its RBF-enabled customers. Likewise, a default rule

forwarding on IP can be used as a stub rule for legacy destinations outside the RBF-enabled AS, and border routers can strip this rule from packets sent from a customer of the RBF-enabled AS to a legacy destination.

ISP are incentivized to upgrade to RBF, since by upgrading to RBF, ISPs gain a network that is fundamentally more secure (because of its default-off nature) while also allowing ISPs to deploy a variety of higher-level network-based services (intrusion detection, caching, transcoders, etc.). This diversified set of services can be made available for customers to use as they see fit, in a flexible and controlled manner. In addition, ISPs can also partner with other businesses, which are not their direct customers, offering them services such as on-path caches or multicast.

8.3 Limitations of RBF

The RBF architecture has several limitations. A first limitation arises when one endpoint changes its network address, while the rule of the other endpoint makes use of that address (see Section 8.3.1). A second limitation is that rules may not be able to express all the forwarding behaviors useful to users (see Section 8.3.2). Lastly, some protocols involving feedback from routers may be more difficult to implement due to RBF's default-off nature (Section 8.3.3).

8.3.1 Using Mobility and DoS Protection Simultaneously

As we have described in Section 3.3, RBF supports host mobility. When a host changes its address from $\texttt{A}_1$ to $\texttt{A}_2$, it can update its rule to send packets to $\texttt{A}_2$ rather than $\texttt{A}_1$. By preserving the same name for the rule, an existing communication will not be disrupted. However, if the other endpoint of the communication, B, uses a rule that drops all packets not sent from the address $\texttt{A}_1$ (*e.g.,* to protect against DoS attacks), packets will not reach B after the change of addresses (as they will be dropped by B's rule). Thus, a communication cannot be maintained while one endpoint changes its network address if the rule of the other endpoint is aimed at protecting against DoS attacks. Next, we briefly discuss how this limitation could be alleviated.

In the simplest instance, the destination could use a different type of rule for the sources that might change their address (information specified by the source). Compared to the other rules used by B to protect against DoS, this rule, $\texttt{R}_X$, should simply not filter packets based on the source address. The destination will be vulnerable to DoS attacks on $\texttt{R}_X$ up to its expiration, and for this reason, $\texttt{R}_X$ should have a shorter lease than the typical rules protecting against DoS attacks. Before $\texttt{R}_X$ expires, if no attack was detected, $\texttt{R}_X$'s lease is renewed. However, if an attack was detected, the destination can stop granting such rules for a period of time as well as reject regular DoS-protection rules to the hosts from which it has received packets on $\texttt{R}_X$ during the attack.

In a more complex solution, but which does not create any vulnerability periods, the destination rule $\texttt{R}_B$ could filter packets based on the name of the return rule $\texttt{R}_A$ (that contains the public key of `A`) rather than the address of `A`. This would prevent hosts with other return rules than $\texttt{R}_A$ to send packets to `B`. However, this approach alone is not enough to protect against DDoS attacks, and we also need to make sure other hosts cannot use the same name as $\texttt{R}_A$ for their rules. To enable its colluders to create return rules with the same rule name as $\texttt{R}_A$, `A` could simply share with them its private key. To prevent this, RCEs should restrict each host to using a single key (or a small number of keys in a predetermined period). Very well coordinated colluders could potentially start many hosts using the same key, but they could attack the destination only once, up to the expiration of the rule $\texttt{R}_B$. `A` could also simply distribute its rule to colluders to be used as return rule, *i.e.,* host `X` would use $\texttt{R}_A$ as its return rule. This can be prevented by the anti-spoofing filters for a restricted set of rules, *e.g.,* unicast rules. Specifically, the anti-spoofing routers can verify whether the source address appears in the return rule. In turn, the destination rule $\texttt{R}_B$ should specify that only such return rules (belonging to the class can be verified by the anti-spoofing machines) are allowed in packets using $\texttt{R}_B$ as the destination rule. We leave the full description of this approach and its implications to future work.

8.3.2 Rule Expressivity

The forwarding behaviors that can be expressed by users through rules are limited in three dimensions.

First, rules cannot modify state at routers, but can only modify state inside the packets. Therefore, users cannot hold state at routers on the path and share this state between packets at the RBF layer (see §7 for a discussion on this topic).

Second, as describes in Section 3, in RBF it is not practical to express forwarding decisions based on any functions other than comparisons of packet and router attributes, *e.g.,* RBF cannot implement forwarding logics based on sum, hash, logarithm of router attributes and addresses.

Third, forwarding patterns requiring variable length fields (such as recording all routers from the path) are also difficult to support in RBF, since the rule forwarding mechanism does not allow adding new attributes during packet forwarding.

These limitations of the rule expressivity were all traded for simplicity, the ability to have provable rule safety and the ease of ensuring policy compliance. The aforementioned limitations of rule expressivity are partly compensated by rules' ability to invoke router-defined functions. The router-defined functions can change the router's state, can implement arbitrary functions and can record variable-sized data in packets, but these functions are controlled by ISPs rather than endpoints. Throughout this dissertation, we showed that a wide range of desirable and previously-proposed forwarding scenarios can be easily expressed using our simple rule syntax.

8.3.3 ICMP-like protocols

Finally, another limitation of RBF is that ICMP-like protocols, in which routers send informative packets back to the source, may not be supported. This limitation is a consequence of RBF's default-off nature, since routers may not be able to send unexpected packets to a host. The ability to deploy such protocols depends on the type of anti-spoofing mechanism used. Specifically, if routers on the path are not be able to send back packets to the source by using the return rule of the packet, such protocols cannot be deployed. Typically, routers could have difficulties sending back packets only for return rules that filter packets based on the source address. Even for such return rules, most anti-spoofing mechanisms, such as ingress filtering or Passport [77] would allow routers on the path to initiate packets back to the source by spoofing the source's address.

Importantly, note that in RBF, for troubleshooting and probing purposes (the typical use of the ICMP-like protocols), endpoints can use router-defined functions deployed by ISPs. For example, router-defined functions can be used to trace routes, record timestamps or perform tracing similar to XTrace [54].

Chapter 9

CloudPolice - Applying RBF to Data Centers and Cloud Computing

RBF describes a general network model that can be applied to other networks than the Internet. In particular, RBF can be applied to data centers and cloud computing environments. Cloud computing is a growing business paradigm in which customers (typically called tenants) temporarily rent machines in data centers owned by cloud computing providers. In essence, cloud computing brings the "pay as you go" model to data centers, allowing clients to dynamically scale up/down to as many machines as needed inside the cloud.

As cloud computing services evolve, they run into analogous security and inflexibility problems as the Internet. Clouds are expanding to very large sizes (currently, hundreds of thousands of machines) and, like the Internet, they have tens of thousands of users. The latter also leads to growing concerns for DoS attacks between users inside the cloud.

In this context, we focus on the problem of providing *network-level access control* between the machines in the cloud, *i.e.,* drop packets undesired by the destination inside the network (close to the source if possible) such that they do not reach their intended destination. Access control is important for security reasons, in order to prevent customers compromising the machines of other customers or mounting DoS attacks on them. Current mechanisms to implement access control policies in clouds are inherited from enterprises, *i.e.,* mechanisms such as VLANs and firewalls. But compared to clouds, enterprise networks have fewer distinct organizations using them, fewer requirements for the flexibility of access control policies, and fewer concerns with internal DoS attacks (for more details see Section 9.1). On the other hand, access control mechanisms designed for the Internet already accommodate all these constraints (large scale, high dynamicity and DoS attacks). In particular, RBF provides access control as one of its properties. Thus, RBF can be applied as previously described in the smaller context of data centers and cloud computing. Moreover, unlike the Internet, each data center or cloud is owned by a single administrative entity and hence it is much easier to adopt new network designs.

However, the same observation (that clouds are owned by a single administrative entity)

leads us to another observation: compared to the Internet, today's data centers are closed environments. In particular, since a single administrative entity owns the entire infrastructure, all the traffic sources in the network can be controlled, monitored and accounted for much better than in the Internet. This observation, leads us to the conclusion that RBF might be simplified when applied to cloud computing.

To this end, we propose CloudPolice, a simplified version of RBF tailored to provide access control in cloud computing. Similarly to RBF, CloudPolice aims to respect the access control policies of the destination and block unwanted traffic close to the source. And, again similarly to RBF, the access control decisions based on destination's policy are made on the control plane and enforced on the data plane by having packets carry a proof of their policy compliance.

On the other hand, CloudPolice takes advantage of the fact that, unlike the Internet, each cloud is owned by a single administrative domain that owns not only the network but also the servers. In particular, cloud computing is typically a virtualized environment, where a trusted layer of software, the hypervisor, sits between each customer machine and the network. In this context, we argue that it is both sufficient and advantageous to implement access control *only* within hypervisors at end-hosts. In this way, CloudPolice has a lower overhead and is easier to adopt. In fact, Cloud providers can adopt CloudPolice through a simple software update to hypervisors, without requiring any hardware upgrades or changes to applications.

Compared to access control techniques inherited from enterprises, CloudPolice can support more sophisticated access control policies, and it can do so in a manner that is simpler, more scalable and more robust.

Next, we describe in more detail the challenges faced by solutions for access control inherited from enterprises when applied to cloud computing and a set of motivating access control policy examples (Section 9.1). We then present an overview of CloudPolice (Section 9.2). Finally, we present more details on the implementation and overhead of CloudPolice (Sections 9.3) as well as on the comparison with other proposals (Section 9.4).

9.1 Background

A major hurdle to the widespread adoption of cloud computing is security, as customers often want to export sensitive data and computation into the cloud. Threats arise not only from privacy leaks to the cloud operating company (outsourcing of data center management is in fact common) but also due to the multi-tenant nature of clouds. For this reason, network-level access control policies are a critical component for preserving security when migrating to the cloud computing model. For example, tenants would want their traffic to be, by default, isolated from all other tenants.

Today, access control in cloud environments is typically provided using techniques such as VLANs and firewalls. These techniques, however, were originally designed for enterprise

environments and as such are ill suited to meet the challenges unique to cloud computing. Specifically, we argue that existing techniques are challenged by three key characteristics of emerging cloud environments:

1. The large scale and high dynamism of cloud infrastructures
2. Diversity in cloud network architectures
3. Multi-tenancy

We expand on each of these challenges in what follows.

Today's clouds house tens of thousands of physical machines, and even more virtual machines that are constantly added and removed (AWS [3] reports over 100 thousand VMs started per day). Current access control mechanisms were not designed to handle such scale and churn. For example, firewalls have problems scaling to large numbers of entries and coordinating access control across multiple firewalls is complex (as described in §9.4), while VLANs do not support dynamic configuration, are limited in scalability and complex to setup and configure [99, 111]. More generally, our observation is that as clouds scale to large numbers of users (AWS reports over 10 thousand users), they will face many of the problems traditionally associated with the public Internet, including DoS attacks between cloud tenants. Such attacks are known to be very difficult to tackle [35, 37, 108, 110] but are not typically the concern of (internal) enterprise access control mechanisms.

Network-diversity is another new challenge for access control. The architecture of data centers has evolved significantly from that of the traditional enterprises and is currently in flux, with many new architectures being proposed, *e.g.,* [33, 59, 62, 63]. These new architectures typically employ multiple paths and require specific routing algorithms and address assignments. Therefore, they severely limit the applicability of current mechanisms such as VLANs and firewalls (as we discuss in §9.4).

Furthermore, multi-tenancy introduces new requirements to access control as intra-cloud communication (*i.e.,* provider-tenant and tenant-tenant) is becoming more popular. For example, Amazon provides their tenants with services such as SimpleDB and Simple Queue Service (SQS); there are also tenants that provide services to other tenants, *e.g.,* mapreduce++ and desktop services [25]. Thus, the intra-cloud communication is likely to require new types of access control policies. Starting from this observation, we next examine a few examples of access control policies that we believe are important for cloud computing environments.

9.1.1 Examples of Cloud Access Control Policies

Our focus for these example policies is on intra-cloud traffic. As we will discuss, several of these policies bring new challenges to cloud policy support and are not supported by existing access control mechanisms or cloud provider APIs. From these examples we will derive a desirable policy model for cloud computing environments (§9.3.1).

Note that the access control policies should have the ability to be defined at a finer granularity than per tenant, *e.g.,* for intra-tenant access control. In this dissertation, we refer to the security principals used in the cloud access control policies as *groups*; this means that a common policy can be defined for all the VMs belonging to the same group and also that policies can be based on (source/destination) groups. A group can be viewed as similar to an AWS security group [3].

Traditional Access Control Policies

The traditional access control lists enable administrators to allow/block traffic through the network based on packet headers. We argue that support for such policies is fundamentally needed in clouds and that new access control mechanisms are required even for implementing these basic access control policies.

Tenant Isolation: The simplest and most common type of access control policy is to block all traffic from other tenants (in particular groups) and this is the default policy in the current cloud environments. Isolation prevents hosts from being compromised and blocks DoS attacks if correctly implemented, *i.e.,* if the traffic from the other tenants is blocked at the source. Note that DoS attacks in the cloud environment can be easier to mount than in the Internet because attackers may not need to compromise hosts to create botnets, but can also simply pay for the hosts used for the attack. Traffic isolation is traditionally implemented by VLANs, which, however, are not a good fit for the cloud environment (§9.4).

Inter-tenant Communication: We expect that the shared environment of cloud computing will enable users to offer each other services more easily than with traditional business models. The main ingredient enabling tenants to more easily offer each other services is the close coupling between the machines in the cloud, *i.e.,* the small latencies and large bandwidths. For example, real time advertising [6, 17] is a fast growing ad-paradigm, projected by Google to serve over 50% of the traffic in the next five years [82]. In real time advertising, a website sends (anonymized) information about a visiting client to a set of ad-providers. The ad-providers send back bids to the website for the ads that should be served to the client. The website then returns to the client the ad of the highest bidder. This ad paradigm requires low latency between advertising providers and providers of web content (which are advertising consumers). Cloud computing offers the perfect environment for such a collaboration between tenants. This example requires the ability to communicate between the ad provider and the ad consumer (for ad bidding, ad retrieval) and to isolate the traffic from the other consumers to avoid DoS attacks.[1] Even this simple communication pattern is not well supported by the traditional enterprise mechanisms such as firewalls and VLANs in the

[1] A mechanism for isolation that also protects against DoS attacks can be very important for services such as bidding and betting, where tenants can indirectly influence each other's bids / bets by creating large drop packet rates for those services.

cloud environment (§9.4).

New Potential Types of Access Control Policies

In addition to the traditional types of access control policies, we believe clouds would likely benefit from new types of access control policies that involve: rate-limiting, fair sharing of the access bandwidth and stateful information on past traffic. Existing mechanisms are (again) not able to support these access control policies.

Fair-Sharing among Tenants: Since multiple tenants may access the services offered by one tenant or by the cloud provider, the entity offering the service may want to implement bandwidth fair-sharing among the groups accessing the service. For example, tenants that have more machines or higher available bandwidth (*e.g.,* are better positioned in the network topology) should not be able to get better service nor impact the services available to other tenants more than their fair share. This is not a feature supported by traditional access control mechanisms, but we believe that it should be supported in cloud computing. Scenarios that show the importance of fair-sharing are storage and database services, *e.g.,* Amazon's SimpleDB [2] and Simple Queue Service (SQS) [1].

Rate Limiting Tenants: As a mechanism, rate-limiting is required by access control to implement the previously mentioned fair-sharing policies. But we argue that in clouds rate-limiting is also important as a policy. For example, in a cloud that charges for bandwidth usage, one tenant `A` may want to rate-limit tenant `B` when accessing `A`'s services. In this case, attackers can financially damage their victims by increasing the bandwidth usage of each VM being attacked; the "pay-as-you-go" pricing model will automatically charge the victim. Moreover, tenants and cloud providers may implement elastic services that automatically add more VMs if the demand increases (*e.g.,* AWS auto scaling). Thus, if there is a DoS attack, more VMs will be added automatically; the charge for the added VMs will fall to the victim.

Allowing Locally Initiated Connections: A common envisioned usage of cloud computing is for virtual desktops [25]. The machines hosting the virtual desktops could use various services also located inside the cloud (to take advantage of the low latencies and large bandwidths), such as SMTP, HR database, EPMAP, a service provided by Facebook, *etc.* Different users with different security credentials can log in the virtual desktop hosts. In these circumstances, it is hard to know in advance what cloud services will be accessed by the virtual desktops. Therefore, it is desired to allow incoming traffic from all of the other groups (of the same tenant or not), but only in response to connections initiated by the virtual desktops. This behavior is typically implemented by stateful firewalls and is not available in current cloud provider APIs.

Time Restricted Access: Tenants offering services may want the other tenants to use these services only during specific time intervals. For example, one tenant could offer backup or archival services, but may be willing to provide these services only during the nighttime in US, when its servers are not heavily serving other clients. Such an access control policy is not typically supported by enterprise mechanisms, and is not currently offered by the cloud provider APIs.

9.1.2 Properties Required for Cloud Access Control

In summary, we believe that the aforementioned challenges call for a new access control mechanism that provides three properties:

1. *Scalability* in handling hundreds of thousands of machines and users.
2. *Network-independence* in decoupling access control from the network topology, routing and addressing.
3. *Flexibility* in providing support for policies in multi-tenant environments such as tenant isolation, fair-sharing, and rate-limiting policies.

9.2 CloudPolice Overview

We now describe CloudPolice, a new access control mechanism implemented in hypervisors, which provides the aforementioned properties. We choose to implement CloudPolice in hypervisors because hypervisors are:

1. *Trusted*. For security reasons, access control needs to be implemented inside a component controlled by hypervisors.
2. *Network-independent*. In this way, our solution is independent of the network architecture and we are avoiding needlessly tying the development of access control to that of specific network equipment and protocols that are known to be slow to change.
3. *Close to VMs*. Thus, unwanted traffic can be blocked before even reaching the network.
4. Have full *software programmability*. For this reason, CloudPolice can provide a broad class of access control polices needed for multi-tenancy.

To achieve scalability, CloudPolice proposes a distributed solution, where hypervisors communicate with each other to exchange access control information. More specifically, CloudPolice uses an approach for enforcing access control on the data plane inspired by Internet mechanisms such as RBF, network capabilities [108, 110] and push-back filters [35, 68]. In CloudPolice, hypervisors of traffic sources attach security credentials to the outgoing

packets, and hypervisors of destination VMs push blocking and rate-limiting filters back to the hypervisors of the sources according to the access control policies of destinations.

We start by discussing the design space for mechanisms by which hypervisors can learn and apply access control policies (Section 9.2.1) and then present the details of our specific mechanism (Section 9.2.2).

9.2.1 Design Space

Although the access control policies are defined by the destination, the policies should be enforced close to the *source* rather than at the destination. This prevents unauthorized traffic from abusing the network (*e.g.,* causing congestion, mounting DoS attacks, *etc.*).

A naïve solution for applying the destination's access control policy at the source is to install all policies and the entire mapping between active VMs and groups in all hypervisors. In this way, the source hypervisor can directly apply the policy of the destination to all the flows sent by its hosted VMs. Unfortunately, this solution scales poorly due to the high churn rate expected for the active VMs. For example, AWS reports about 100 thousand new VMs started per day; in a data center with 100 thousand servers [3, 21, 58], this translates into more than 100 thousand update messages sent per second on average (peak update rates would likely be much higher).

Another extreme solution is to distribute no policies to hypervisors, but use a centralized repository for policies and group membership. Hypervisors then consult this repository for each new flow and possibly cache the access control policies. However, the centralized resolution service is likely to represent a tempting target for DoS attacks. Moreover, the centralized service has to sustain very high availability and low response times. For example, assuming an average of 10 new flows per second per server [97], and a 100K server cloud, the centralized service would need to process 1M flows per second on average (again, with higher peak rates expected). Note also that caching may be ineffective since the traffic is expected to be randomly distributed [73, 97] and since policies and VM locations change in time.

For these reasons, we propose a *distributed* solution. In CloudPolice, hypervisors need only know the policies of their hosted VMs and not the policies of any other group in the cloud, nor the group membership. In order to learn which flows should be allowed, blocked or rate limited, CloudPolice uses a *runtime* approach in which hypervisors communicate with one another using a secure channel as we describe next.

9.2.2 CloudPolice's Design

Fig. 9.1 shows a high level description of the operation of CloudPolice.

When a new flow is initiated by a VM, the source hypervisor sends a *control packet* containing the security credentials of the source VM (step 1). In particular, the control packet contains the security group to which the source VM belongs to; this packet is sent

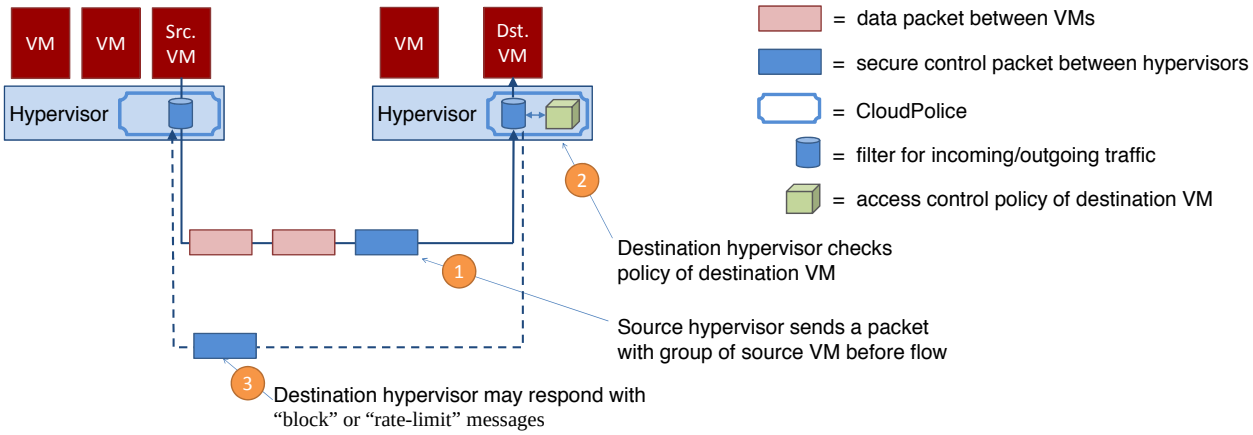


Figure 9.1: CloudPolice Overview

before the the packets belonging to the flow. When the destination hypervisor receives such a control packet, it checks the policy of the group of the destination VM (step 2). If the policy allows the traffic, the destination hypervisor creates state for this flow; subsequent packets will be forwarded up to the destination VM by using this entry. If the traffic is not allowed or should be rate limited, the destination hypervisor will send a control packet back to the source hypervisor to block or rate-limit the flow or the VM (step 3). By default, VMs are blocked if the policy contains no rule for that traffic.

Note that the control packets containing the source's security credentials can be seen as similar to how each RBF packet contains a proof of its policy compliance. However, unlike RBF, the security credentials are contained in a separate packet for each flow (rather than in each packet) and do not use cryptographic techniques. These simplifications are meant to make the scheme more easily deployable and are specific to data center environments. In data centers, one can guarantee that each misbehaving host can be blocked by its hypervisor and also that control packets cannot be injected by malicious VMs (for more details see the security discussion in Section 9.3.4).

9.3 Detailed Design and Implementation

We now present more details on the design and implementation of CloudPolice.

9.3.1 Proposed Policy Model

We aim to design a policy model that supports the examples presented earlier in Section 9.1.1. Table 9.1 presents the syntax of one such policy model. The policy is formed by a sequence of entries processed in order. Each entry is of the form `if` *`condition`* `then` *`action`*. For

```
Policy := [Entry]+
Entry := if Condition then Action

Condition := Predicate [(and | or) Predicate]*
Predicate := Field Op (value | ANY) | locally_initiated
Op := > | < | ≤ | ≥ | == | ≠
Field := source_group | HeaderField | crt_time | StatisticsField
HeaderField := src_addr | src_port | protocol | dst_port | ...  (others)
StatisticsField := StatisticsType period (vm-level | group-level)
StatisticsType := in_bytes | out_bytes | in_flows | out_flows | rej_flows

Action := block | allow | rate-limit RateType
RateType :  (absolute value | relative weight) (vm-level | group-level)
```

Table 9.1: CloudPolice's access control policy model for incoming traffic.

every new incoming flow, CloudPolice iterates over all entries and checks the condition of each entry. CloudPolice applies the action corresponding to the first matching condition.

Conditions are designed to capture three new types of access control necessary for clouds — (i) group-based, (ii) time-based, and (iii) usage-based. Group-based access control allows applying the same action over a group of VMs, *e.g.,* all VMs that belong to a tenant. Time-based access control allows tenants to control traffic at specific periods of time. Usage-based access control allows tenants to control traffic based on statistics such as the number of active connections or traffic volume. For example, if we want to prevent port scanning, we can keep a count on how many ports (or how many rejected flows) have been used by a group of VMs within a given period of time, and block the traffic if the count exceeds a threshold.

Table 9.1 presents the specific syntax of the policy model. To simplify the explanation, we assume the access control policy belongs to a given group `A`. *src_group* represents the group of the remote (source) VM trying to contact the VM belonging to group `A` at which the policy is applied. *crt_time* represents the current GMT time; for ease of use, the syntax could expose specific parts of the current time like day, month, hour, *etc.* To enable the use of the current time in the access control policy, we assume servers are using a form of time synchronization such as NTP. *locally_initiated* is a predicate that is true if the current flow (for which the policy is applied) has been initiated by the VM belonging to group `A`, *i.e.,* the received traffic for which the policy is applied is only in response to this locally-initiated traffic. `HeaderField` represent fields in the packet header, while `StatisticsType` encodes different types of statistics gathered in a given period such as the number of incoming flows (in_flows), rejected flows (ref_flows), *etc. vm-level* statistics are maintained at each VM belonging to group `A` while *group-level* statistics represent an aggregate value for all the VMs belonging to `A`. Rate limiting at *vm-level* means that the rate-limit value is applied

independently at each machine belonging to group `A`, while rate-limiting at the *group-level* means that the rate-limit value specified in the policy is an aggregate at the level of all the machines in group `A`. For example, if groups `B` and `C` are rate-limited at the group level by `A`'s policy, the entire aggregated traffic from `B` and `C` to `A` is rate-limited at the value specified in the policy.[2] For a better understanding, please also see the examples below.

Examples

We now present a few examples in order to illustrate how the above policy model is able to express the desired access control policies in clouds. We assume the following examples represent the access control policy of group `A`.

- Tenant isolation: Group `A` blocks all traffic except that generated from its own group.
  `if` *source_group* $\neq$ `A then` *block*
- Providing services: Group `A` provides a service to group `B` on port 80.
  `if` *source_group* $==$ `B and` *dst_port* `==` 80 `then` *allow*
- Rate limiting: Rate limit traffic incoming at any machine belonging to group `A` from the set of all machines belonging to group `B` to at most $100Mbps$.
  `if` *source_group* $==$ `B then rate-limit` *absolute* $100Mbps$ *vm-level*
- Timed access: Tenant `A` provides a backup service to tenant `B` but wants `B` to be able to access it only during nighttime (within a given time zone, we assume GMT here).
  `if` *source_group* $==$ `B and (`*crt_time.hour* $>$ 22 `or` *crt_time.hour* $<$ 6`) then` *allow*
- Traffic size limit: Tenant `A` wants to limit the total traffic received from group `B` to $1GB$ per day.
  `if` *source_group* $==$ `B and` *in_bytes* $<$ $1GB$ *day* `then` *allow*
- Fair sharing: Group `A` would like to provide equal service throughput in case of congestion to groups `B` and `C`. The following policy assigns them equal relative weights.
  `if` *source_group* $==$ `B then` *rate-limit relative* 1 *group-level*
  `if` *source_group* $==$ `C then` *rate-limit relative* 1 *group-level*

Policy restrictions

We now discuss how policies should be restricted to ensure that different entries of the policy do not conflict with each other, causing inconsistent semantics. The described policy model is more prone to conflicts than traditional access control policies since it is more complex.

First, we require that the *locally_initiated* predicate, which has semantics only at the destination VM, is not used with the *group-level* rate-limiting action, which has semantics at the level of the entire destination group.

The second restriction is that the policy does not contain multiple entries with rate-limiting actions defined on overlapping flow sets, *i.e.,* that the set of conditions that can

[2]Note that the access control policy does not specify how the rates should be distributed between the different source or destination VMs, but only that the entire traffic is rate-limited.

lead to rate-limiting actions have void intersections. For example, if one entry rate-limits the total traffic from port range 50-100 to $20Mbps$, while another entry rate-limits the traffic from group `B` to $10Mbps$, the semantics of the rate-limit value for the traffic sent by group `B` using port range 50-100 are not clear. One can simply apply the first matching entry and restrict the flow along with the rest of the flows that are matched by the same entry, however, we impose this restriction for the clarity of policies.

Note that there can be other policies with difficult to understand semantics. For example, when some tenants are rate-limited to an absolute value and others using a relative value, the latter are in fact sharing the remaining bandwidth of the former. In another example, some tenants could be rate-limited at the entire group level while others at the VM-level, which results in the relation between the service allocation between the different tenants dependent on the communication pattern. However, we leave further restrictions that could ease the use of the policy model to future work

Limitations of current implementation

Our implementation currently supports a limited subset of the described policy model in two ways. First, our current implementation does not support the *group-level* keyword, *i.e.,* it does not support group-aggregated state, nor rate-limiting policies at the group level. Second, our implementation does not handle rate-limiting across an aggregate set of flows, *i.e.,* we only support rate-limiting of individual flows or entire VMs. We plan to address these limitations in future work (see Section 10).

In the rest of this section we describe a detailed design an implementation with the aforementioned limitations.

9.3.2 Design Details

CloudPolice requires maintaining state at hypervisors as well as communicating policy information and actions between hypervisors using special control packets. Next, we discuss these in turn.

Soft State: CloudPolice maintains soft state (*i.e.,* removed after expiration) to enforce the policy actions (block, remove and rate-limit). After the expiration of the soft state, the entire process for setting up the state is restarted. Soft state makes it easier to support VM migration and handle packet losses. In our current design, revocation (of state and policies) is also handled through the expiration of the soft state. However, explicit state invalidation on policy updates could be implemented, by using control packets between hypervisors, in a similar fashion with the rest of CloudPolice's design.

Control Packets: There are three types of control packets sent between hypervisors:

1. `source`: This type of control packets is sent by a source hypervisor to inform the

destination about the security credentials of the source. The `source` control packet is presented in step 1 in Fig. 9.1.

2. `action`: This type of control packet is sent by a destination hypervisor to block or rate limit the traffic of a specific source. The `action` control packets are presented in step 3 in Fig. 9.1.

3. `query`: Finally, the `query` control packets are sent by a destination hypervisor to query the source hypervisor about the security credentials of a specific VM (*i.e.,* requesting the hypervisor of the source VM to resend the `source` control packet). As we shall describe, the `query` control packet is used when the original `source` control packet is lost.

The `source` control packet contains the group to which the source address belongs to. When receiving a `source` control packet, hypervisors insert the mapping between the network address of the source VM and its group into an internal cache. The `action` control packet specifies the type of action that should be applied, which can be one of (a) *block* or (b) *rate-limit*. The `action` control packet also contains the granularity at which this action should be applied. The granularity can be per flow or per VM in our current implementation. Rate-limiting actions also contain the value of the rate to which the respective flow/VM should be limited.

In our current design, control packets are distinguished from the rest of the data packets by using a special transport protocol number in the IP header, *e.g.,* the protocol number 254, which is reserved for testing. Note that CloudPolice is independent of the mechanism used to distinguish control packets from data packets and, in the future, other approaches are possible. Also note that only hypervisors can send control packets. The control packets sent by VMs or incoming to the data center from outside Internet hosts are dropped.

Lost/Reordered Control Packets: First, assume the original `source` control packet is lost. If a destination hypervisor receives a flow for which it has no entry, it sets up a querying state for the flow and sends a `query` control message to the source hypervisor.[3] At the receipt of the `query` control message, the source hypervisor resends the `source` control packet. There is a timer associated with the querying state and a new `query` request is made to the source hypervisor when that timer expires, *i.e.,* in the case the `query` control packet is lost.

Second, assume an `action` control packet, sent by the destination hypervisor, is lost and the destination keeps receiving unwanted traffic. After sending each `action` packet to block traffic, the destination hypervisor sets up a short term state to block the incoming traffic, waiting for the `action` message to arrive at the source hypervisor. If packets are still received after the short timeout expires, the destination hypervisor resends the `action` packet. In

[3]Note that the rest of the incoming packets belonging to this flow may be buffered into a small sized queue for performance reasons.

case of rate limiting, the destination monitors the incoming rate and, if the limit is not respected, it sends back a new `action` control packet to the source hypervisor. However, rate-limiting control packets are usually sent periodically to the source due to policies based on traffic statistics (which are updated periodically) or policies at the group-level (which need periodic refreshing).

Note that the above timers used to retransmit lost control packets can be set proportional to the maximum latency of the network.

Policy Distribution and Updates: We envision two models for distributing and updating policies to hypervisors. In the first model, the cloud provider uploads the group policy to the hypervisor at the VM startup and updates it at all the group members when the policy changes. Since policy changes should be infrequent, we do not expect this service to be a burden for the cloud provider. Moreover, cloud providers must have a service for starting and shutting down VMs; we expect that installing and updating policies can be easily piggybacked on top of this existing service.

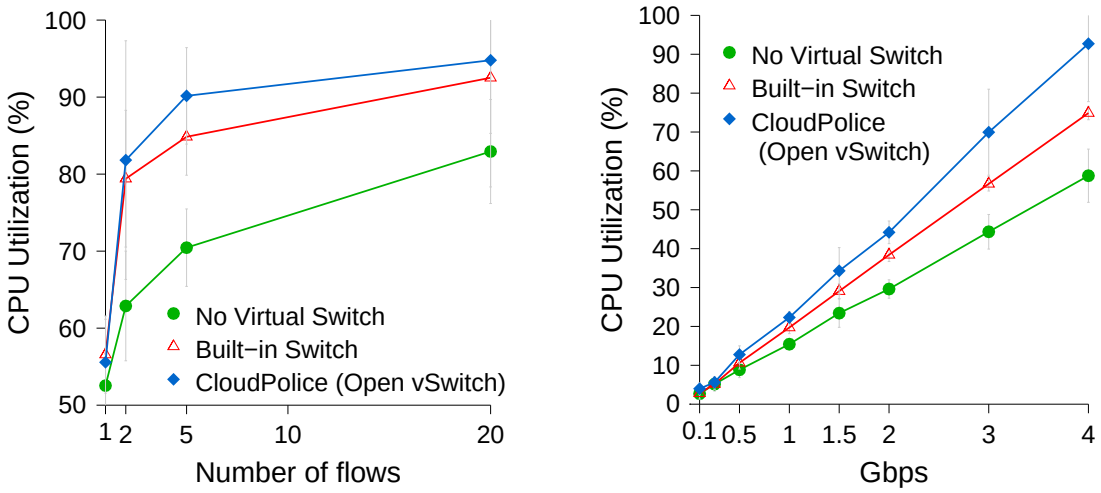
In a second model, VMs can directly communicate their policies to hypervisors. This model does not require a policy management service from the cloud provider but requires and additional API in both the hypervisor and the VMs.

Stateful Policies: As we have described, some policies can be based on past history of traffic (*i.e.,* traffic statistics). For each entry of the policy that uses traffic statistics and that has been activated by one or several of the incoming flows, we start a thread that collects the desired statistics. The tread is waken periodically, using the period(s) specified in the entry of the policy. The policy is re-applied at the end of each period for the active flows matching the respective policy entry.

The policies containing the *group-level* keyword require state maintained at the granularity of a group. Such policies require either direct communication between the hypervisors hosting one group's VMs or the use of a centralized state repository offered by the provider, to which all the hypervisors hosting VMs of that group communicate. As mentioned, we do not discuss these options in this dissertation.

9.3.3 Implementation and Evaluation of Overhead

We have implemented CloudPolice using Open vSwitch [88]. More specifically, we have prototyped CloudPolice as a NOX module [100] that controls a single Open vSwitch instance and is co-located with it on the same physical machine. Thus, we control the virtual switch of the hypervisor with a controller that runs as well in the (user space of the) hypervisor. We use this approach for the simplicity of the implementation (Open vSwitch already has control hooks implemented through OpenFlow [23]) and since we have observed that the performance is not significantly affected compared to running Open vSwitch alone. A future and more efficient implementation might be directly performed inside Open vSwitch.



(a) Multiple TCP flows

(b) Single UDP flow

Figure 9.2: Overhead of CloudPolice in terms of utilization of a single CPU core compared to using no virtual switch and to using the default (built-in) switch of the hypervisor.

As described earlier, we use soft state and timeouts to invalidate the actions specified by the CloudPolice policy.

In this dissertation, we only evaluate the overhead added by our implementation of CloudPolice to the end hosts. We use a setup formed by two machines running Xen connected through a 10Gbps cable. Each machine is a dual-socket server with four 2.8GHz Intel Xeon (X5560) cores per socket.

Due to extra processing of packets and of policies, CloudPolice adds extra latency and incurs additional CPU utilization at the end hosts compared to other types of access control mechanisms such as VLANs and in-network firewalls (which incur other types of costs). We measure the overhead of CloudPolice compared to two different baselines. (1) The common case of using a virtual switch inside hypervisors (as it seems to be performed by today's cloud providers). (2) We assume that VMs can access the network interface card directly, bypassing the virtual switch, *e.g.,* see [14].

Measuring overhead in virtualized environments is complex, because the overhead resides in many components: guest domain (DomU), driver domain (Dom0) and the virtual machine monitor (Xen in our case) [80, 92]. Since our setup does not support accessing the network interface directly from the VM (the second baseline above), we assume the overhead in the guest domain to be similar for CloudPolice and the two aforementioned baselines, and we restrict to measuring the overhead in the driver domain. The purpose of this evaluation is not to measure overhead associated with a specific virtualization environment, but instead, provide a relative comparison among different forwarding schemes.

Latency: There are two types of latency that can be added by CloudPolice: to all packets and to the flow setup (*i.e.,* to the first packet only). The virtual switch in the hypervisor adds a negligible latency to each packet – less than 0.1ms per round-trip. CloudPolice adds essentially no additional latency compared to that, assuming that the state allowing the traffic has been instantiated in the virtual switch (*i.e.,* the flow has been set up).

For the flow setup, in our measurements, CloudPolice, adds $\approx$ 1.5ms round-trip latency. Note that this additional latency of CloudPolice is mostly an artifact of our prototype implementation, done as a NOX controller in userspace. We expect a production implementation to reduce this latency significantly.

Therefore, in practice, CloudPolice adds no noticeable overhead in terms of flow completion time.

CPU overhead: Typically, the additional CPU usage incurred by CloudPolice on top of the virtual switch is negligible, since CloudPolice is not acting on a per-packet basis. However, different virtual switches have different overheads. Fig. 9.2 shows the CPU overhead of CloudPolice (built on top of Open vSwitch) compared to the built in bridge module of Xen and compared to not using any virtual switch. Since CloudPolice has an analogous CPU utilization as Open vSwitch (not displayed in the charts), we believe that the additional overhead of CloudPolice compared to the switch built in the hypervisor can be reduced in the future through improved implementations (*e.g.,* either by implementing CloudPolice in a different virtual switch module or by improving the Open vSwitch module). Note that we use a simple allow/deny access control policy and we leave the evaluation of the policy engine to future work.

Fig. 9.2(a) shows the CPU utilization when using multiple concurrent TCP flows that send as much traffic as possible. As mentioned, we start these flows from the driver domain (Dom0). The CPU utilization is scaled to a single core, *i.e.,* 50% means half of one CPU core is used and the others are idle (remember that our testing machine has 8 cores). Starting from around two TCP flows we fully utilize the 10Gbps link (the throughput is the same for the different plotted series). As one can see, the price paid for using a virtual switch compared to using no virtual switch is at most 20% of one CPU core for $10Gbps$ of traffic. For example, assume we divide the machine to four VMs, each with two CPU cores. If each VM is sending at $2.5Gbps$[4] the added overhead associated with using the virtual switch for each VM relative to its CPU power is 2.5% (5% of a single CPU core for each VM and the VM has two cores). Since, I/O intensive applications are typically not concurrently CPU intensive, this small overhead should not be noticed by users. More generally, the use of a virtual switch allows us to implement access control and achieve all the advantages described for CloudPolice (scalability, network independence, flexibility), and thus we believe that this overhead is a modest price to pay.

Fig. 9.2(b) shows the CPU utilization when using a single variable rate UDP flow (again

[4]In our tests, VMs with two cores can communicate with more than $2.5Gbps$.

sent from Dom0). We believe the higher CPU utilization compared to the case of TCP is due to the hardware and kernel support for TCP in the network stack, and due to the inefficiency of the traffic generator/sink when manipulating UDP packets. In this case, at $4Gbps$ the largest overhead due to using a virtual switch is about 30% of one CPU. For our previous example with four VMs running on the server, the overhead for each VM when sending $1Gbps$ of UDP traffic is less than 4%.

9.3.4 Security Analysis

We now analyze possible attacks originated inside clouds by either malicious tenants or compromised VMs. We consider three classes of attacks: (1) circumvent the access control policies and reach a destination VM with unauthorized traffic, (2) mount a DoS attack using unauthorized traffic, and (3) mount DoS attacks by using authorized traffic. Next we discuss how CloudPolice addresses each of them.

Circumvent CloudPolice Policies

There are two potential cases in which a VM could receive undesired packets that circumvent its policies.

The first case is when the hosting hypervisor is compromised. However, when a hypervisor is compromised, access control is not the main concern. For example, a compromised hypervisor might directly corrupt and spoof on the private data of the VMs that it hosts.

The second case is when a hypervisor receives fake information about the sender. This could occur even without compromised hypervisors, *e.g.,* if VMs could inject spurious control packets into the network. To prevent this case, CloudPolice requires hypervisors and ingress routers to drop control packets sent by VMs or incoming to the cloud from external sources. Therefore, only hypervisors can send control packets. For this reason, CloudPolice does not require packet encryption or other security protocols to protect control packets, and thus can generate control packets with low overhead.

Above, we have assumed hypervisors are not compromised. However, if hypervisors are compromised, they could send spurious control packets. Next, we discuss mechanisms to protect against the case when some of the hypervisors are compromised.

Compromised Hypervisors: If at least one hypervisor is compromised, the attacker can inject arbitrary control packets into the network. By inserting fake `source` control packets, destinations can be fooled into accepting packets from unauthorized sources. These packets can be used to corrupt the destination. By inserting spurious `action` control packets, other VMs can be blocked and not allowed to communicate.

To prevent these attacks, CloudPolice needs extra support from the network. The simplest solution is to implement an *anti-spoofing* mechanism. This prevents corrupted hypervisors from sending packets with spoofed IP addresses, which do not belong to any of the

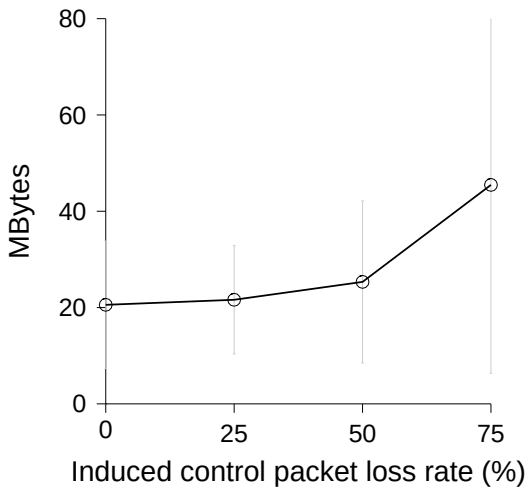


Figure 9.3: Unwanted traffic received by recipient. We simulate large drop rates for control packets.

VMs running on it. Thus, anti-spoofing prevents corrupted hypervisors from lying about the credentials of any VM in the network except for the VMs located on that machine. It also prevents a corrupted hypervisor `H` from blocking a communication that does not involve any of the VMs running on top of `H`. Note that cloud providers today already implement anti-spoofing mechanisms to prevent hosts inside the cloud to be used to mount DoS attacks on hosts external to the cloud [4].

The cloud provider can further prevent a compromised hypervisor `H` to lie even about the credentials of the VMs running on top of it. For this purpose, in conjunction with anti-spoofing, the cloud provider must sign the binding between the VM's IP address and its group. In this way, each `source` control packet contains the binding between $< source\_ip$, $source\_group >$ signed with the private key of the cloud provider. The destination hypervisor uses the public key of the cloud provider to verify whether these credentials are correct. Therefore, even compromised hypervisors cannot send malicious `source` control packets containing fake groups, and cannot circumvent the access control policies.

DoS with unauthorized traffic

In order to not add additional latency to the flow setup, CloudPolice optimistically allows all flows and requires destination hypervisors to block unwanted traffic through `action` control packets sent back to the source hypervisor. Thus, some amount of unauthorized traffic can be injected into the network. Attackers can try to take advantage of this and attempt to DoS a victim VM `V` by sending it unauthorized traffic. Since, as we shall see, only a small

amount of unauthorized traffic is allowed through the network, attackers can attempt to increase the amount of unauthorized traffic in the network in two ways.

First, an attacker VM `X` can try to prevent control packets from reaching its hypervisor $\texttt{H}_X$. In this way, $\texttt{H}_X$ would not know that `X`'s traffic should be blocked. To mount this attack, `X`'s group can flood $\texttt{H}_X$ (possibly with authorized traffic), and in this way induce losses of control packets. A simple fix for this attack is to prioritize control packets in switches (most switches today support QoS). However, even in the absence of traffic prioritization, with a loss rate of 50%, only two control messages are required to block one VM.

Fig. 9.3 quantifies the previous attack by showing the amount of unwanted traffic received at a destination from a source blasting at 3Gbps with UDP packets. We use the same two-machine setup described in Section 9.3.3. To simulate the worst case scenario for the cloud provider, we assumed control plane traffic is not prioritized and competes directly with data plane traffic. Thus, we simulated significant loss rates for control packets. The amount of unwanted traffic received in case of lost control packets depends on the retransmission timeout for control packets; we use a retransmission timeout of 50ms, which we believe can accommodate even large data centers.[5] The amount of unwanted traffic received also depends on the round-trip time (RTT) between the source and the destination. In our setup, the RTT is very small (less than 1ms); for larger networks, the values in Fig. 9.3 should be increased with the amount of traffic sent in the additional RTT. For example, if the RTT is 10ms larger, we should add 3.75MB for each data point in Fig. 9.3 and also add the same amount for each lost control packet. However, we leave further investigation of larger RTTs to future work. We note that the received traffic shown in Fig. 9.3 is higher than the traffic sent during the RTT even when there are (almost) no control packets lost, amounting to ≈50ms of traffic. We believe this additional traffic is due to (1) buffering at endpoints, buffers that are sent by the virtual switch despite the blocking signal and (2) the fact that the blocking state is not instantiated immediately in the virtual switch. For example, the virtual switch may not update the same flow entry immediately (an entry allowing the flow is always inserted at the flow start), the software path installing the blocking state competes for a lock on the flow table with the forwarded packets, *etc.* We believe these shortcomings can be addressed by improved future implementations.

Importantly, the amount of traffic shown in Fig. 9.3 is received periodically, where each period equals the timeout of the soft state associated with blocking a VM. In our implementation, we use a value of 20 seconds for this period. Hence, even with a very high loss rate of 50%, to fill up 3Gbps of traffic at a single victim VM, the attacker would need ≈300 VMs, each of them being able to send 3Gbps.

In a second attack, `X` tries to exhaust the filters available at hypervisors (either at $\texttt{H}_X$ or at `V`'s hypervisor $\texttt{H}_V$). In the unlikely event for this to occur (since we expect CloudPolice to store a small amount of memory and memory to be abundant), CloudPolice can aggregate

[5]We use this timeout for both the retransmission of a `block` control packet, as well as for the `query` control packet sent in case the `source` packet was lost.

the per-flow state into per-VM state and block the VMs that have a significant fraction of their flows blocked.

Note that compromised hypervisors can send any amount of unauthorized traffic. However, if the rate of compromised hypervisors is that high to mount a DoS attack, DoS is in fact not the main worry of the cloud provider.

DoS with authorized traffic

Access control policies ensure that only authorized traffic competes for bandwidth. However, access control does not protect against floods of authorized traffic between colluders that have access to a shared link with the victim (a well known limitation of network capabilities [110]). Due to the virtualized environment, this situation is much more common in clouds than in the Internet. For example, an attacker can attempt to DoS a VM `V` by sending a lot of authorized traffic to VM `X` located on the same physical machine with `V`.

Preventing DoS attacks in the above scenario requires *performance isolation* between VMs, in addition to access control. Performance isolation can be implemented using an orthogonal mechanism, and we do not discuss it in this dissertation.

However, we point out that CloudPolice's mechanism could be used to implement performance isolation by rate limiting VMs in case of congestion with the same mechanism used for access control. For example, if the hypervisor runs N VMs, the VMs sending traffic to each one of these N VMs can be rate-limited together to $1/N$ (with equal shares between them or a more sophisticated distribution algorithm). In other words, the bandwidth is evenly shared between the *destination VMs*. This approach does not guarantee the fair sharing of a congested link, but does guarantee fair sharing of the bandwidth at the destination. In this way, a performance isolation framework based on CloudPolice can prevent DoS attacks with authorized traffic as well. For instance, a legitimate VM cannot be DoSed with authorized traffic sent to another VM located on the same physical machine. Existing approaches to implement performance isolation do not block DoS attacks with authorized traffic because the sharing is based on the *source VMs* rather than the destination ones. For example, one solution for performance isolation is to statically share the bandwidth, *i.e.,* if there are N VMs on each machine, each can send up to $1/N$ of the available bandwidth. This approach cannot prevent DoS attacks since the attacker can use an arbitrary number of VMs to send traffic. In another example, (dynamic) fair sharing per source VM has been recently proposed to be implemented in clouds [95]. This approach is more effective in mitigating DoS attacks compared to static bandwidth sharing, but still suffers from the problem that the attacker can use more source VMs than the legitimate traffic, and thus reduce the bandwidth of the destination VMs proportionally.

9.4 Related Work

A traditional mechanism to implement access control is through the use of VLANs. However, VLANs have several limitations. First, since VLANs couple access control and switching, they cannot be applied to new network topologies such as FatTrees [33], BCube [62], DCell [63], *etc.*, nor to topologies that use L3 routing instead of switching. In addition, VLANs have much overhead both in spanning tree creation/maintenance and in switching between VLANs (which typically requires L3 routing) [56,99]. VLANs are also limited by the number of hosts in one VLAN and the number of VLANs in a network [111]. Furthermore, VLANs do not offer the flexible policies proposed in this dissertation.

Firewalls can also be used to block unwanted traffic at the source, *e.g.,* by placing them at the first-hop switches. However, this approach presents a major maintenance overhead, since every time a destination changes its policy, all the firewalls at all possible sources need to be updated. Moreover, in order to support group-based policies, firewalls either need to create an entry for each VM in the group, or group the VMs by the same IP prefixes and create an entry for each prefix. Neither of the solutions is desirable because the former faces a scalability limit, and the latter makes VM address management unnecessarily complex.

Centralized controllers such as OpenFlow and Ethane [23, 43] can be used to provide access control. However, these approaches are network-dependent, *i.e.,* they require changes to the switching hardware. Open vSwitch [88] can be used to achieve network-independence, but it still requires a centralized controller. Using Open vSwitch with a centralized controller inherits all the drawbacks of centralized approaches – the centralized controller could be a scaling bottleneck, and a potential attraction for DoS attacks from the tenants. Moreover, the OpenFlow API implemented by Open vSwitch is designed for switches, but much richer policies can (and should) be implemented in hypervisors. VL2 [59] also discusses a mechanism to make address assignment independent of the underlying topology, but still makes use of a centralized service for access control. Thus, it suffers from the same drawbacks listed above for Open vSwitch.

AWS [3] offers a limited set of access control policies [4] such as isolation and on/off access between groups. We are not aware of public information describing AWS's implementation.

Chapter 10

Future Work

In this section, we list a few avenues for future work for RBF and CloudPolice.

10.1 Extending RBF

Automated Tools for Rule Creation and Policy Compliance Verification: The deployability of RBF depends, in addition to incentives for ISPs, on the ease of use for end users. We envision most end users to not have to do anything else compared to today. Specifically, rules will be created for end users by their ISPs, enterprises (for employees), by third parties (*e.g.,* middlebox service providers) or by firewall-like tools located on the end user's machines. We envision all these entities will use automated tools to create rules. For example, such tools could be similar to the current firewall configuration programs and create simple rules given various properties desired by end users. An end user could select for which ports to receive data, could chose to enable mobility or could decide to use a middlebox service such as intrusion detection. Note that middlebox rules can be created by programs located on the end user's machine or can be created by the middlebox service provider (*e.g.,* the user fills out an application form to use the middlebox service on the website of the provider of the service). Thus, the existence of automated tools to create rules is fundamental for the ability to deploy RBF and is an important place for future work.

In addition to the tools for creating rules, providers and end-users should also use automated tools for verifying the policy compliance. Typically, these tools should verify simple properties of the rules, such as which senders are allowed to send and which other destinations are named in the rule (as discussed in §4).

High-speed RBF Router: In this dissertation we have presented a software router prototype to evaluate the performance of forwarding on rules. RBF is also compatible with high-end routers that use specialized hardware, rather than the general-purpose x86 hardware we have used. Increasingly, high-end routers rely on network processors [49,103] for data

plane processing (*e.g.,* Cisco's high end CRS routers [11] use network processors). We expect network processors can accommodate the flexibility RBF requires in per-packet processing, and we leave building such a prototype to future work.

RBF is also a candidate for implementations based on "hybrid" forwarding planes as proposed by Casado *et al.* [44]; In brief, the proposal in [44] is to use TCAMs to cache recent forwarding decisions. Incoming packets that hit in the cache are forwarded at high-speed using TCAM lookups while the remaining packets are forwarded using in-software processing. The software processing creates entries in the TCAM for future packets. Since rules are deterministic and likely shared across multiple packets, we expect RBF would achieve the high cache hit rates required by a router design similar to the one in [44].

Extending RBF's Design: In Section 8.3 we have describe a set of limitations of RBF. While we have hinted to possible solutions to address the respective limitations, we leave a full integration of such approaches into the architecture to future work.

We also leave to future work a full extension of RBF to enable ISPs, in addition to sources and destinations, to benefit from the flexibility offered by RBF. ISPs can encapsulate packets and use their own rules (*i.e.,* a form of tunneling), however, one might imagine integrating such support into the rule forwarding mechanism in the future.

10.2 Deploying RBF Through HTTP

Fueled by the explosive growth of video traffic and HTTP infrastructure (e.g., CDNs, web caches), HTTP is fast becoming an integral part of the Internet architecture. Starting from this observation, we argue that HTTP may be a fertile ground (more so than IP) for deploying solutions and techniques provided by the clean slate proposals. There are three reasons for this. First, HTTP is more extensible than IP in that is easier to add functionality to an HTTP server or proxy than an IP router. Second, HTTP allows incremental deployability, as clients are able to explicitly pick the the proxies and servers implementing the desired functionality. Third, HTTP already provides a massive infrastructure, thus alleviating a major hurdle a clean-slate architecture needs to overcome.

Many of the ideas proposed by RBF can be implemented through HTTP. In particular, the HTTP header can contain an RBF rule. This would improve the current Internet in two ways.

First, it would enable end users to have better control over HTTP proxies. Through instructions in the HTTP header rule, clients and servers could configure forward and reverse proxies, as well as specify which functionality should or should not be invoked for different requests. For example, endpoints could redirect packets across a set of HTTP proxies to avoid failed routes (*e.g.,* similar to RON [34]), and activate intrusion prevention to detect malware for some of the requests.

Second, RBF can also be used to provide access control for one endpoint `A`, assuming

that all the communication of `A` is enforced to occur through (a) an HTTP proxy or through (b) a set of indirection servers that relay HTTP requests between endpoints as proposed in [89]. Specifically, the HTTP proxy / relay server can verify whether the endpoints trying to contact `A` have valid rules to grant them access as discussed in this dissertation. In addition, `A` could also use rules to push packet filters and firewall rules at these proxies / relaying servers.

Note that having rules in the HTTP header also inherits RBF's policy compliance property. Specifically, the ability of the end users to use the respective proxies and invoke their functionalities is contained by the rule's policy compliance mechanism. This should reduce CDN's and ISP's concerns for exposing a larger API at their HTTP proxies and servers.

We leave a comprehensive description of the detailed components necessary to deploy RBF through HTTP to future work.

10.3 Extending CloudPolice

There are two directions we are actively pursuing to extend CloudPolice as presented in this dissertation.

First, we are working on implementing global policies that require state maintained at the level of an entire security group. For example, to rate-limit the traffic between two groups, we need to maintain aggregated state about the rates experienced by all the machines in both groups. This extension requires a mechanism to share state between the hypervisors hosting VMs belonging to the same group, as well as distributed algorithms for how to use this state to implement the access control policies.

Second, we are investigating a class of dynamic policies that are controlled at runtime by VMs based on application-level semantics. For example, source VMs could obtain access capabilities dynamically (*e.g.,* obtained through external websites), allowing them to access certain groups. Destinations, on the other hand, could block unwanted traffic at runtime, based on application-level semantics.

Chapter 11

Conclusion

In this dissertation we introduce the idea of forwarding rules and develop the design of a rule-based forwarding (RBF) architecture. A network based on RBF is flexible and, at the same time, policy compliant. To demonstrate flexibility, we consider a wide variety of forwarding scenarios and show how rules can be used to implement them, *e.g.,* forwarding through waypoints, middleboxes, using router extensions, achieving mobility, *etc.* We show that policy compliance is a required architectural property in order to (a) prevent the additional flexibility available to end users to enable new malicious attacks and (b) be able to prevent DoS attacks. Through a prototype implementation and evaluation of RBF-related overheads we show that RBF is feasible for implementation at Internet scale.

We also show how to apply the concepts behind RBF to cloud computing. We describe CloudPolice, an access control mechanism implemented in hypervisors that is scalable, network independent and provides flexible policies.

We hope that the ideas developed in this dissertation will help the computer science community to arrive at better network designs in the future, as the importance of networks in the human society continues to grow.

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Appendix A

Static Analysis of Rules for Forwarding Loops

We describe an algorithm called *Loop-Free Rule (LFR)* that signals whether a rule can result in a forwarding loop. A rule can result in a forwarding loop when the packet is sent from one destination to another indefinitely (or through a sequence of destinations); dropping such a packet requires the forwarding layer to rely on the TTL field in the IP header or other similar mechanisms. [1]

Importantly, this static analysis is required due to our design decision to keep the rule forwarding mechanism as simple as possible. Indeed, forwarding loops could be avoided by using runtime mechanisms. In one example, all the destinations in a rule can be encoded in the packet using a loop-free data structure such as a Directed Acyclic Graph (DAG); the forwarding mechanism can impose the restriction that the destination address attribute can only be set such as to advance in this data structure (*i.e.,* can only set the destination to children of the current destination in the DAG). Note that such a mechanism is not restricting rule expressivity (all rules that cannot generate forwarding loops can be expressed this way), but it incurs some extra overhead for storing the loop-less data structure in the packet.

Rules have to be loop-free regardless of the attributes inserted by senders. However, we are not concerned with the fact that attributes can be changed outside the rule forwarding engine by malicious routers or by functionality invocations; such changes could lead to loops at the overlay level rather than at the forwarding level, *i.e.,* malicious routers have to perpetually change the attributes to preserve the loop. Indeed, changing the packet attributes by a malicious router is equivalent to the insertion of random attributes by the sender, and the packet has to be loop-free unless it is changed again by another malicious router.

We next present the *LFR* algorithm that detects those rules that can result in forwarding loops. *LFR* is based on the creation and analysis of the rule finite state machine (FSM), which captures the states in which the packet can be during forwarding – we defined the

[1] Note that obviously we are assuming there are no routing loops, which are outside the scope of RBF.

rule FSM in the next section. *LFR* relies on the following observation: a forwarding loop is possible only if the rule FSM has a cycle that involves states with multiple destination addresses.

Note that in RBF all attributes have unsigned integer values.

A.1 Rule FSM

We assume a rule R with packet attributes pa_1,...,pa_n using router attributes ra_1,...,ra_m. The possible sets of values for these attributes are PA_1,...,PA_n and respectively RA_1,...,RA_m. We note the concatenation of all the packet attributes with $pa = pa_1...pa_n$ and $PA = PA_1 \times PA_2 \times ...PA_n$. Similarly, $RA = RA_1 \times ...RA_m$. Note that PA_i and RA_i are finite sets whose size depends on the binary representation of the attributes; in our current implementation most attributes are represented using two byte values. Without loss of generality, we commonly assume in this section that these values represent positive integers.

We define the following finite state machine: $(\Sigma, S, S0, \delta, F)$. The input alphabet is $\Sigma = RA$. The set of states $S = (PA \times Invoke) \cup \{drop\}$. The last state is a special state where the packet has been dropped. *Invoke* can be seen as equivalent to a special attribute encoding whether the packet is currently *in process* at a higher layer functionality or not. The $Invoke$ attribute equals the value of the invoked functionality in R. Specifically, $Invoke = \{invoke_i | invoke_i \in R\} \cup \{NO\_INVOKE\}$, where $invoke_i$ is the identifier of a router defined function invoked by R, and NO_INVOKE is a special value for all other cases when the packet is simply in the process of being forwarded (which is the common case).

By $R(s, r) = s_r$ we note the application of the rule R when the packet is in state $s \in S$, the router attributes at the current router are $r \in RA$, and the resulting state of the packet is $s_r \in S$. Note that it does not make sense to apply the rule to the *drop* state (since the packet is dropped). The $Invoke$ attribute is ignored at the application of the rule, but its value is set as the invoked functionality identifier, in case a router function was invoked, or to NO_INVOKE otherwise. Also, note that some of the state transitions occur irrespective of the router attribute values; these are similar to the FSM λ transitions.

In this dissertation, we are interested in the properties of the FSM regardless of the initial state or of the final states, and thus S_0 and F are not of interest for our analysis (all states can be considered both as initial and final).

Lemma 1. *A forwarding loop is possible in a rule R* iff. *R's FSM contains a cycle and at least two of the states from the cycle have different values for the destination address packet attribute.*

Proof. First, assume the rule can lead to a forwarding loop. This means that the packet destination address attribute has to change during the packet's forwarding. Since the space of all packet attribute values is finite (PA), this implies there is a cycle in the FSM.

Second, if such a cycle in the FSM exists, this means that the rule could oscillate between multiple destinations indefinitely if the router attribute values match specific requirements. □

The problem with the rule FSM is that the number of states ca be very large. Next, we present aggregated states, a method used by *LFR* to reduce the number of states.

Aggregated States: We define an *aggregated state* s_{agg} as a subset of $PA \times Invoke$, $s_{agg} \subset PA \times Invoke$, *i.e.,* for an aggregated state, the packet attribute values are generalized to have multiple values (such as entire intervals).

We generalize the rule application for aggregated states as such: $\forall s_{agg} \subset PA \times Invoke$, $R(s_{agg}, r) = \cup_{s \in s_{agg}}(R(s, r))$. In other words, $R(s_{agg}, r) = \{S_o\}$, $s.t.\ \forall s \in s_{agg}$, then, $R(s, r) \in S_o$ and $\forall s_o \in S_o$, $\exists s \in s_{agg}$ s.t. $R(s, r) = s_o$.

Similarly to above, we also extend the rule application notation to subsets of the router attribute space. If $R_s \in RA$, we define $R(s_{agg}, R_s) = \cup_{r \in R_s}(R(s_{agg}, r))$

We note with $S_p \subset S$, the subset of all possible states after the rule application, $S_p = R(PA \times Invoke, RA)$, *i.e.,* we obtain all possible states at the end of the rule if we apply the rule to the aggregated state representing all the possible inputs. In other words, S_p contains all the possible packet states before the packet is forwarded to IP or to upper layers; in the latter case, the *invoke* attribute captures the identifier of the invoked functionality. Note that not all possible states are in S_p, *i.e.,* in general $S \neq S_p$. For instance, if an attribute `X` is always set to the value 3 in the rule, all states where `X` $\neq 3$ are not in S_p.

A.2 The *Loop-Free Rule (LFR)* algorithm

The high level idea is to reduce the states of the rule FSM to a set of aggregated states (such that the analysis is feasible), but, at the same time preserve a large enough number of states to permit an accurate detection of one rule's ability to create forwarding loops.

LFR has three steps. The first step identifies a set of k aggregated states $S\_FSM_{agg} = \{s1_{agg}, ...sk_{agg}\}$ such that $s1_{agg} \cup s2_{agg} \cup sk_{agg} = S_p$ and $s1_{agg} \cap s2_{agg} \cap sk_{agg} = \Phi$. These are the states of the reduced FSM. The second step creates all the possible transitions between these states. Note that we are not interested in the inputs to the FSM that make these transitions possible (*i.e.,* the router attributes) since we want to prevent loops regardless of the value of the router attributes, which are under router control. Finally, the third step takes the decision of whether the rule can lead to a forwarding loop or not based on whether the FSM graph created has cycles or not.

The first two steps of *LFR* rely on a procedure that computes $R(s_{agg}, RA)$, *i.e.,* the set of states that can result by applying the rule R to any input state in s_{agg} and using any value for the router attributes. This procedure is called *Rule Application Procedure (RAP)* and we present it next.

RAP: Rule Application Procedure: *RAP* simulates the rule application and in fact performs a simple form of symbolic execution (simple because rules have a very constrained format). *RAP* identifies an aggregated state as the *"current state"* in which the packet is during the rule application. For conditional branches, this analysis is replicated and the simulation continues for each branch. The current state is initialized with an input value s_{start} and is modified during the rule application, when some of the attributes are set to specific values or when conditional branches are taken.

The input of *RAP* is an aggregated state s_{start} that corresponds to the states the packet can be in at the beginning of the rule application.

The output of *RAP* is a set of aggregated states, $S\_OUT_{agg} = \{s1_{agg}, ..., sk_{agg}\}$. There are two important properties of the resulting aggregated states. First, $s1_{agg} \cup s2_{agg} \cup sk_{agg} = R(s_{start}, RA)$. Second, these states correspond to different `sendto` and `invoke` actions in the rule, and thus, they capture the essential properties for detecting loops.

Below, we note with "`*`" an attribute not set, *i.e.,* can have any value.

RAP defined by the following procedures to be applied when the corresponding rule actions are encountered:

1. `if(packet.a <op> value)`: If the current state can accommodate both decision branches, we *replicate* the state, one for each branch, and continue to apply *RAP* for each of the resulting states. We set the value of the `a` attribute in each resulting state to an interval or value conforming to the decision taken. For example, if the initial value of attribute `a` is `*` and `a` takes positive integer values, then for the condition `if(packet.a < 5)` the two created states will have intervals $0 - 4$ and $5 - max$ for `a`'s value.

2. `if(router.a <op> b)`: The state is replicated as before, but remains unchanged for both branches, regardless of `b`, which can be a packet attribute or a constant.

3. `if(packet.a <op> packet.b)`: The state is replicated and the generalized values for `a` and `b` are possibly constrained if one of them is constrained; *e.g.,* if `<op>` is "<", `b` is constrained to the range $2 - 7$ and `a` is unconstrained (*i.e.,* `*`), the two resulting states will constrain `a` to $0 - 6$ on one branch and $3 - max$ on the other.

4. `packet.a = value`: The `a` attribute is set in the current state to that specific value.

5. `packet.a = router.b`: `a`'s value in the current state is set to `*`.[2]

6. `packet.a = packet.b`: `a` takes the generalized value of `b`.

[2]If we assume that a router attribute's value is the same throughout the rule execution, an improvement can be made to this algorithm, to also constrain the value of the packet attribute, if previous constraints have already been made on the router attribute's value. For simplicity, here we make no such assumption.

7. `sendto D`: Set `packet.destination` to `D`, set the `Invoke` attribute to NO_INVOKE and add the aggregated state to S_OUT_{agg}.

8. `drop` (or the end of the rule): Discard current aggregated state and add `drop` to S_OUT_{agg} (if not already in).

9. `invoke func_id`: Set the `Invoke` attribute to $func\_id$ and add the current state to S_OUT_{agg}.

RAP terminates when all states resulted from different branch replication are added to S_OUT_{agg}.

Lemma 2. *$R(s_{start}, RA) = s1_{agg} \cup s2_{agg} \cup ...sk_{agg}$, where si_{agg} states are computed as the result of applying RAP to s_{start},* i.e., *$RAP(s_{start})$. This means that RAP simulates all possible rule executions.*

Proof. This lemma is proven by the construction of RAP and by exhaustively taking into consideration all the possible statements in rules. After each rule action or decision, the set of aggregated states maintained by RAP covers all the possible states resulted from the application of that action/decision.

Explaining this argument a bit more, assume there is $s_x \in R(s_{start}, RA)$ but $s_x \notin s1_{agg} \cup ...sk_{agg}$. This means $\exists s_0 \in s_{start}$, $\exists r_0 \in RA$ s.t. $R(s_0, r_0) = s_x$. Since we apply the rule to $s_{start} \ni s_0$, this situation can only occur if: (1) the presented algorithm does not take a possible rule branch or (2) the generalized state at the end of the branch of s_x does not contain s_x. First, by branching, RAP covers all possible branches, by definition setting the value of all attributes to the most general value that the attribute can have on that branch regardless of the value of the router attributes. Second, within each branch, for each possible rule action, it is easy to see from the algorithmic RAP steps above that the value of any attribute is set to the most general subset that can result from that action. Thus, $R(s_{start}) \subset s1_{agg} \cup s2_{agg} \cup ...sk_{agg}$.

On the other hand, assume that exists state $s_y \in s1_{agg} \cup ...sk_{agg}$ but $s_y \notin R(s_{start}, RA)$. This means RAP maintained as current states, some states that are not actually possible. By analyzing all the rule actions and decisions exhaustively captured by RAP, one can see that such a situation cannot occur unless the router attributes or the initial packet attributes have constrained value ranges; however, we cannot make such an assumption in the presence of malicious adversaries and arbitrary semantics associated to router and packet attributes. □

First Step of *LFR*: The first step is to apply the RAP algorithm to the most general aggregated input state $PA \times Invoke$, by computing $RAP(PA \times Invoke)$, *i.e.,* all attributes have the value $*$. This can be seen as computing $R(PA \times Invoke, RA)$, *i.e.,* all the possible results of the rule applications when the values of the packet attributes and of the router attributes are not constrained.

The result of the first step of *LFR* is a set of aggregated states S_FSM_{agg}. The elements of S_FSM_{agg} represent the aggregated states of the finite state machine we will use to detect loops.

S_FSM_{agg} is initialized with the output S_OUT_{agg} of $RAP(PA \times Invoke)$. *LFR* makes two further (simplifying) adjustments to this set. First, if $\exists i, j$ s.t. $sj_{agg} \subset si_{agg}$, we only retain si_{agg} in S_FSM_{agg}. Second, if $\exists i, j$ s.t. the destination attribute in sj_{agg} is equal to that in si_{agg} and $si_{agg} \cap sj_{agg} \neq \Phi$ and both aggregated states have the *invoke* attribute equal to NO_INVOKE, we remove si_{agg} and sj_{agg} from S_FSM_{agg} and insert $si_{agg} \cup sj_{agg}$. These adjustments do not pose security threats since the destination address attribute (relevant in loop detection) is not aggregated.

Lemma 3. $S_p = s1_{agg} \cup s2_{agg} \cup ...sk_{agg}$, *where* $si_{agg} \in S\_FSM_{agg}$, i.e., *are the aggregated states computed by the first step of LFR.*

Proof. This simply results from Lemma 2 and the definition of S_p, $S_p = R(PA \times Invoke, RA)$. The two minor adjustments described above do not remove states nor introduce new states to S_FSM_{agg}. □

Note here that a malicious router could create states that are not in S_p (*e.g.,* set random values to the packet attributes) but, as mentioned, this is equivalent to having the sender insert random attribute values; the state of the packet after the next (non malicious) router will get back in S_p.

Lemma 4. $s1_{agg} \cap s2_{agg} \cap ...sk_{agg} = \Phi$, *where* $si_{agg} \in S\_FSM_{agg}$.

Proof. By construction, all states with the *invoke* attribute set to NO_INVOKE are disjoint either by having different destination address attributes or by the fact that we join overlapping aggregated states with the same destination address (see above adjustments). All the other states have different *invoke* attribute values (for simplicity, we assume here that the functionalities invoked in the rule are distinct, *i.e.,* no functionality is invoked by two different *invoke* statements).[3]. This shows that there are no states that belong to multiple aggregated states identified by the first step of *LFR*. □

Second Step of *LFR*: We now create the transitions between the states in S_FSM_{agg} identified at the previous step. For this purpose we use the following algorithm:

for all $si_{agg} \in S\_FSM_{agg}$ **do**
  compute $R(si_{agg}, RA)$ // i.e. apply $RAP(si_{agg})$
  for all $sj_{agg} \in S\_FSM_{agg}$ **do**
    if $R(si_{agg}, RA) \cap sj_{agg} \neq \oslash$ **then**
      add link from si_{agg} to sj_{agg}

[3]The algorithm can easily be extended when this is not the case (*e.g.,* by uniquifying the identifiers).

end if
end for
end for

In other words, the second step of *LFR* applies the *RAP* algorithm for each of the aggregated states si_{agg} in S_FSM_{agg} (identified at the first step). If some of the states resulted belong to other aggregated states sj_{agg} from S_FSM_{agg} then we create an edge from si_{agg} to sj_{agg}. Such a link means that it is possible for the packet to go from a state in si_{agg} to a state in sj_{agg} for some values of the router attributes; given the above construction of the aggregated states, this signifies that the destination address or the invoked functionality changed.

Third Step of *LFR*: In this step, *LFR* decides whether the rule R can lead to forwarding loops.

At the end of the second step, we obtain a graph of aggregated states; the nodes in this graph are the members of the S_FSM_{agg} set and are the result of the first step of *LFR*, while the edges are added in the second step. We name this graph R_FSM, since it simulates an (aggregated) finite state machine of the rule R.

We further remove all the "invocation" states in R_FSM, *i.e.,* those states that have the *invoke* attribute different than NO_INVOKE, since these are not relevant for forwarding loop detection. When removing such an aggregated state si_{agg}, we connect each of the neighboring states sj_{agg} that have transitions to si_{agg} to each neighboring state sk_{agg} reachable from si_{agg}, unless sj_{agg} and sk_{agg} are already connected through a direct edge.

If the resulting graph has cycles, *LFR* returns `TRUE`, meaning that the rule can lead to cycles, otherwise *LFR* returns `FALSE`.

A.3 Proving Rule Safety

Theorem Forwarding-loop Safety. *A rule R is guaranteed to not result in forwarding loops if the* LFR *algorithm returns* `FALSE`*.*

Proof. A network loop requires at least two distinct destination address attributes. The packet finds itself at each point during its life in exactly one of the states in S_FSM_{agg}; this is proven by Lemma 3 and Lemma 4. All the possible transitions between the aggregated states in S_FSM_{agg} are captured by the transitions created at the second step of the *LFR* algorithm; this is true since the *RAP* algorithm captures the rule application, see Lemma 2. (Note that the removal of states with the *invoke* attribute different than NO_INVOKE at the third step of the *LFR* algorithm does not remove any possible path between two of the remaining states since we add extra edges between all neighbors.) Finally, by construction, each state has a single value for the packet destination attribute.

Therefore, a forwarding loop actually requires a cycle in the graph constructed by the

LFR algorithm. Thus, if the *LFR* algorithm returns `FALSE`, the rule cannot lead to forwarding loops regardless of the value of the router attributes encountered on its path.

□

Although we do not present it here, the *LFR* algorithm does not create false positives, *i.e.,* cases when *LFR* returns `TRUE` but the rule cannot lead to a forwarding loop.

A.4 Complexity

We make the following notations: RS represents the rule size, measured in the number of statements: actions (`sendto` and `invoke`), settings of attribute values and branching conditions in the rule; D is the number of destinations in the rule (the number of distinct targets of `sendto` actions); and A is the number of packet attributes of rule R.

The complexity of the first step of the *LFR* algorithm is $O(RS)$. (We ignore here the complexity of maintaining the generalized attribute values.) This represents the complexity of applying the *RAP* procedure. The complexity of the second step is: $O(|S\_FSM_{agg}| \cdot (RS + (|S\_FSM_{agg}| \cdot A))$. To understand this result, see that the *RAP* algorithm is applied for each aggregated state resulted from the first step; the result of each such application is also compared with all the other states to check for intersections; finally, this check depends on the number of packet attributes in the rule. The number of aggregated states ($|S\_FSM_{agg}|$) is on the order of the number of `sendto` actions in the rule. Therefore, the complexity of the second step is $O(D \cdot RS + D^2 \cdot A))$. The complexity of the last step of the *LFR* algorithm is, in the worst case, $O((|S\_FSM_{agg}| + |E|) \cdot |S\_FSM_{agg}|)$, where $|E|$ is the number of edges between the aggregated states in S_FSM_{agg}. This step tries to detect whether there are cycles in the R_FSM graph and we apply DFS from each of the aggregated states. Hence, the last step of the *LFR* algorithm has a worst case complexity of $O(D^3)$.

In conclusion, the complexity of the *LFR* algorithm is $O(D \cdot RS + D^2 \cdot A + D^3)$. If we consider RS to be dominated by D^2 (RS also contains the `sendto` actions) and A to be on the same order as D, we obtain a simplified worse case complexity of $O(D^3)$.

Since rules are bounded to a small size (the rule description is bounded to 256 bytes in our current implementation), we expect all rules to be easily statically analyzable; *e.g.,* we expect the number of destinations D to be well less than a dozen. For example, with our current implementation, one could fit 24 destinations in a 256B rule encoding using an anycast-like rule (note that RS for such a rule has a value around 50).

Appendix B

Static Analysis of Rules for Local Loops

A local loop is a continuous invocation of functionality at a router; the packet invokes a router functionality, then it returns to the forwarding layer after the functionality invocation ends, but the rule invokes the functionality again creating a loop. Such a loop could also be created by looping among multiple functionalities offered by the same router.

Similarly to forwarding loops, local loops can be prevented through the use of runtime mechanisms. For example, invocation statements can be associated an order number in the rule; at each host, the forwarding layer remembers which functionality returned the packet back to it and prevents the rule from accessing another functionality with the same order number or a smaller one. This approach does not constrain the rule expressivity, since the invoked functionalities at a router can always be consistently ordered, but it does incur extra overhead for packet forwarding and packet payload (to store the ordered of the invocation statements).[1]

To detect the rules that can lead to local loops, one could use the same LFR algorithm as for detecting forwarding loops presented in Appendix A. The difference is in the condition to be checked on the R_FSM graph.

However, the assumption used in Appendix A that router attributes can have arbitrary values at each rule application is sometimes too constraining for this type of analysis. More precisely, since local loops occur at a single router, we can assume that the router attributes have the same value before and after the functionality invocation. For example, the router address will not change between each functionality invocation. This relaxation of the algorithm comes in handy for functionalities that are accessed over and over again at multiple routers such as the multicast registration example in Section 3.3. We expect that router attributes have the same value before and after the functionality invocation, because we expect the duration of the packet processing at the invoked functionality to be much shorter

[1]This approach would not allow invoking a function twice at the same router.

than the period between router attribute updates.[2]

For this reason, we modify slightly the algorithm presented in Appendix A. Each packet attribute pa_i can take values in $PA_i \cup RA\_ID$ (instead of only PA_i), where RA_ID is a set containing the identifiers for each router attribute used by the rule R. This extends the symbolic execution of the RAP algorithm in the following way: the value of each packet attribute can also signify that the respective attribute has been set to the value of a router attribute. To achieve this change, the RAP algorithm presented in Appendix A is modified in the following case:

5. `packet.a = router.b`: `a`'s value is set to the identifier of the attribute `b` $\in RA\_ID$.

We construct an algorithm called *Local Loop-Free Rule (LLFR)* to detect whether a rule can result in a local loop. As the LFR algorithm presented in the Appendix A, $LLFR$ has three steps; the first two steps are identical to those of LFR (but using the modified RAP algorithm as described above). Thus, at the end of these two steps the R_FSM graph is computed.

The third step of the $LLFR$ algorithm removes all the states in the R_FSM graph that have the `invoke` attribute different than NO_INVOKE; the edges adjacent to these states are also removed. Note that these states are not relevant for local loops since the packet is being forwarded between different network nodes.

$LLFR$ returns `FALSE` (*i.e.,* the rule cannot result in a local loop) if the remaining graph has no cycles, otherwise returns `TRUE`.

Theorem Local-loop Safety. *A rule R is guaranteed to not result in local loops if the $LLFR$ algorithm returns* `FALSE`.

Proof. The argument is similar to the forwarding loops, so we rely on Lemmas 2, 3, and a similar Lemma to 4 from Appendix A.

Assume the rule forms a local loop. This means that after each functionality invocation, the rule directly uses another functionality invocation. Since the number of functionalities is limited, and since each state of the reduced FSM created by our algorithm are disjoint, representing distinct functionalities, this implies a cycle between one or more aggregated states has to exist. But such cycle is detected by the $LLFR$ algorithm. Note that for a local loop, all the aggregated states in the cycle have to invoke functionality (*i.e.,* the `invoke` attribute be different than NO_INVOKE) and so, we can safely ignore the other (forwarding) states as described above.

□

[2]Even if this were not the case, loops could hardly be created because the rule creator would have to guess the new value of the router attribute.