ShortMAC: Efficient Data-Plane Fault Localization*

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Abstract

The rising demand for high-quality online services requires reliable packet delivery at the network layer. Data-plane fault localization is recognized as a promising means to this end, since it enables a source node to localize faulty links, find a fault-free path, and enforce contractual obligations among network nodes. Existing fault localization protocols cannot achieve a practical tradeoff between security and efficiency and they require unacceptably long detection delays, and require monitored flows to be impractically long-lived. In this paper, we propose an efficient fault localization protocol called $ShortMAC$*, which leverages* probabilistic *packet authentication and achieves 100 – 10000 times lower detection delay and overhead than related work. We theoretically derive a* lower-bound guarantee *on data-plane packet delivery in* $ShortMAC$*, implement a* $ShortMAC$ *prototype, and evaluate its effectiveness using the SSFNet simulator and Linux/Click routers. Our implementation and evaluation results show that* $ShortMAC$ *causes* negligible *throughput and latency costs while retaining a high level of security.*

1 Introduction

Performance-sensitive services, such as cloud computing, and mission-critical networks, such as the military and ISP networks, require high assurance of network data delivery. However, *real-world* incidents [2, 7, 9, 23, 32, 52] and studies [10, 16, 43, 59] reveal the existence of compromised routers in ISP and enterprise networks, and demonstrate that current networks are surprisingly vulnerable to data-plane attacks: a compromised router or a dishonest transit ISP can easily drop, delay, inject or modify packets on the forwarding path to mount Denial-of-Service, surveillance, man-in-the-middle attacks, etc. Unfortunately, current networks do not provide any assurance of data delivery in *adversarial* environments, and lack a reliable way to identify misbehaving routers that jeopardize packet delivery. For example, a malicious or misconfigured router can "correctly" respond to `ping` or `traceroute` probes while corrupting other packets.

Though end-to-end path monitoring [13, 22] and multipath routing [20, 21, 31, 44, 54, 55, 57] can mitigate data-plane attacks to some extent, they are proven to render poor performance guarantees [47, 59]; without the exact knowledge of which link is faulty, a source node would need to explore an *exponential* number of paths in the number of faulty links in the worst case. As illustrated in Figure 1 where the default route from S to D is path $(1', 2, 3', 4)$, end-to-end monitoring only indicates if the current *path* is faulty without localizing a specific faulty link (if any) of a compromised or misconfigured router on the path. In the worst case, S needs to explore 2^4 paths to find the path with no faulty links, i.e., path $(1, 2, 3, 4)$.

Therefore, data-plane *fault localization* has been widely recognized as a promising remedy for securing data delivery [10, 11, 16, 59]. In a nutshell, fault localization enables a source to monitor data forwarding at each hop and localize abnormally high packet loss, injection, and/or forgery on a certain *link*. Such information about link quality can be utilized for two vital purposes. First, by excluding detected poor links the source can select high-performance routing paths to carry its traffic, thus eliminating the expo-

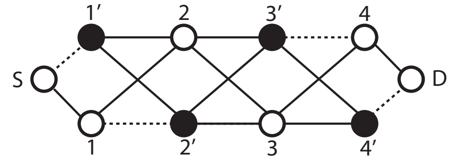

Figure 1. Exponential path exploration problem for end-to-end monitoring. Dotted links are faulty links of malicious routers (black nodes).

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nential path exploration problem as depicted in Figure 1. Second, fault localization provides **forwarding accountability** which proves to be a *necessary* component for enforcing contractual obligations between participating nodes in a contractual networking service such as the Internet or wireless mesh networks, as demonstrated by Laskowski and Chuang [33].

Unfortunately, existing fault localization protocols suffer from *security*, *efficiency*, and *agility* challenges in the presence of strong adversaries. (i) **Security and efficiency:** Sophisticated attacks such as framing and collusion attacks and natural packet loss tend to break fault localization protocols (e.g., Fatih [43], ODSBR [14], Watchers [18], AudIt [10], Network Confessional [11], etc) or lead to heavy-weight protocols (to prevent sophisticated attacks). (ii) **Agility:** In addition, recent *secure* and *relatively* lightweight protocols [16, 59] leverage *packet sampling* or *flow fingerprinting* to prevent packet modification attacks while reducing communication overhead. However, in addition to high storage overhead, these techniques result in long detection delays and require monitored paths to be long-lived (e.g., after monitoring 10^8 packets over the same path in *Statistical FL* by Barak et al. [16]), which is impractical for networks with short-lived flows and agile routing paths.

In this paper, we propose ShortMAC, an efficient fault localization protocol to provide a *theoretically proven* guarantee on end-to-end *data-plane* packet delivery even in the presence of sophisticated adversaries. More specifically, we aim to guarantee that, *given a correct routing infrastructure, a benign source node can quickly find a non-faulty path along which a very high fraction of packets can be correctly delivered*. Our key insights are two-fold:

Insight 1. We first observe that localizing data-plane faults along a communication path can be reduced to monitoring packet *count* (number of received packets) and packet *content* (payload of received packets) at each router on that path. Furthermore, if packets can be *efficiently* authenticated, packet count also becomes a verifiable measure of packet content, because forged packets (with invalid contents) will be dropped by the routers and manifest an observable deviation in the packet count. Thus, routers can dramatically reduce storage overhead by storing counters instead of packet contents.

Insight 2. We also observe that we can achieve a high packet delivery guarantee via fault localization by *limiting* the amount of malicious packet drops/modifications, instead of perfectly detecting *each single* malicious activity. Furthermore, strong *per-packet* authentication to achieve perfect detection of *every single* bogus packet is unnecessary for *limiting* the adversary's ability to modify/inject bogus packets. Instead, the source can use much shorter packet-dependent random integrity bits as a weak authenticator for each packet such that each forged packet has a

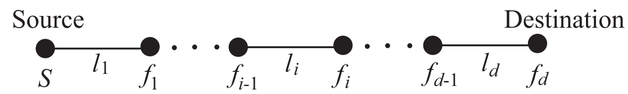


Figure 2. An example path and notation.

non-trivial probability to be detected. In this way, if a malicious node modifies or injects more than a threshold number of (e.g., tens of) packets, the malicious activity will cause a detectable deviation on the counter values maintained at different routers. Essentially, ShortMAC traps an attacker into a dilemma: if the attacker inflicts damage worse than a threshold, it will be detected, which may lead to removal from the network; otherwise, the damage is limited and thus a guarantee on data-plane packet delivery is achieved.

Contributions. 1) We propose a data-plane fault localization protocol ShortMAC that achieves high security assurance with 100 - 10000 times lower detection delay and storage overhead than related work.
2) We derive a provable lower bound on successful end-to-end packet forwarding rate, by limiting adversarial activities instead of perfectly detecting every single malicious action which would incur high protocol overhead.
3) We theoretically derive the performance bounds of ShortMAC and evaluate ShortMAC via SSFNet-based [6] simulation and Linux/Click router implementation.

2 Problem Statement and Setting

We consider a general multi-hop network model where *routers* relay packets between *sources* and *destinations*, such as the ISP and enterprise networks. Throughout the paper, we follow the notation as illustrated in Figure 2. We denote the routers in a path by $f_1, f_2, \ldots, f_{d-1}$, the destination by f_d, and the link between f_{i-1} and f_i by l_i.

2.1 Adversary Model

The goal of an adversary who controls malicious routers is to sabotage data delivery at the forwarding path. Instead of considering an individual forwarding attack, we seek a general way of defining malicious forwarding behavior. We identify packet dropping and packet injection as the two fundamental data-plane threats, while other data-plane attacks can be reduced to these two threats as follows: (i) packet modification is equivalent to dropping the original packet and injecting a fabricated packet, (ii) packet replay can be regarded as repeated packet injection, (iii) packet delay can be treated as dropping the original packet and later injecting it, and (iv) packet misrouting can be regarded as dropping packets along the original path and injecting them to the new path. A formal definition follows:

Definition 1 An $(x, y)-$***Malicious Router*** is a router that intentionally drops *up to* a fraction *x* of the legitimate data packets from a source S to a destination f_d, and injects *up to* y spurious packets to f_d, pretending that the packets originate from S. The *misbehavior space* of such a malicious router comprises (i) dropping packets, (ii) injecting packets on any of its adjacent links which we call **malicious links** (non-malicious links are called **benign links**), (iii) strategically claiming arbitrary local state (e.g., number of packets received) to its own advantage, or (iv) colluding with other malicious routers to perform the above attacks.

Such a strong attacker model is not merely out of theoretical curiosities, but has been widely witnessed in practice. For example, outsider attackers have leveraged social engineering, phishing [7], and exploration of router software vulnerabilities [2, 9] and weak passwords [23] to compromise ISP and enterprise routers [52]. Also, in a 2010 worldwide security survey [1], 61% of network operators ranked infrastructure outages due to misconfigured routers, which also fall under our attacker model, as the No. 2 security threat.

Furthermore, we assume that an adversary knows the cryptographic keys of controlled routers, and can eavesdrop and perform traffic analysis anywhere in the network. The protocol parameters are public; as a consequence, the adversary may attempt to bias the measurement results to evade detection or frame honest links. However, the adversary cannot control the natural packet loss rate on the links in the path, because this would constitute a physical-layer attack which can be dealt with through physical-layer protections. We consider attackers with polynomially bounded computational power which cannot break cryptographic schemes, e.g., encryption or Message Authentication Codes (MAC).

2.2 Problem Statement

Our paper focuses on providing data-plane fault localization for a lower-bound guarantee on data-plane packet delivery. In this section, we define communication epochs, detection thresholds, faulty links, and finally we formalize fault localization.

Definition 2 An end-to-end communication is composed of a set of consecutive **epoch**s. An epoch *for an end-to-end path* is defined as the duration of transmitting a sequence of N data packets by a source S toward a destination f_d along that path. The epochs are *asynchronous* among different paths.

The introduction of epochs facilitates detection and formal analysis as we show later.

Definition 3 Given a **drop detection threshold** T_{dr} (i.e., fraction of dropped packets) and an **injection detection threshold** T_{in} (i.e., number of injected packets), a link l_i is defined as **faulty** iff: (i) more than T_{dr} fraction of packets are dropped on l_i by f_i in an epoch, *or* (ii) more than T_{in} packets are injected by f_i in an epoch, *or* (iii) the *adjacent* router f_i or f_{i+1} makes l_i appear faulty in an epoch.

When T_{dr} and T_{in} are carefully set based on the prior knowledge such that the natural packet loss and corruption are below T_{dr} and T_{in}, respectively, a faulty link must be a malicious link.

Definition 4 $(N, \delta)-$**Data-plane Fault Localization** is achieved iff: given an end-to-end communication path p, after a **detection delay** of sending N packets, the source node S of path p can identify a specific faulty link along that path (if any) with false positive or negative rate less than δ.

Definition 5 $(\Omega, \theta)-$**Guaranteed Forwarding Correctness (Guaranteed Data-Plane Packet Delivery)** is achieved iff: after exploring at most Ω paths, a source can find a **non-faulty path** (*if any*) along which all routers have correctly forwarded at least θ fraction of the source's data packets sent along the path to f_d.

To achieve a guaranteed θ, we need to *bound* (not necessarily *eliminate*) the adversary's ability to drop packets and to inject packets so that if the adversary drops more than α percent of packets or injects β bogus packets, it will be detected with a high probability. A formal definition follows.

Definition 6 For an epoch with a sufficiently large number of data packets by a source, $(\alpha, \beta)_\delta-$**Forwarding Security** is achieved iff two conditions are simultaneously satisfied:

1. (*Low False Negative Rate*) When the adversary drops more than α percent of the data packets on a single link, or injects more than β fake packets on a single link, the source will detect at least one of the malicious links under the adversary's control with probability at least $1 - \delta$;

2. (*Low False Positive Rate*) The probability of falsely incriminating at least one benign link is at most δ.

2.3 Scope and Assumptions

Since we focus on data-plane security at the network layer, we assume the following network control-plane and link-layer mechanisms, each of which represents a separate line of research orthogonal to ours. (i) We can borrow existing secure routing protocols [24, 27, 45, 58] by which nodes can learn the genuine network topology and the source can know the outgoing path. (ii) We assume secure neighbor identification so that a node upon receiving a packet knows which neighbor sends that packet, which can be achieved via link-layer authentication. (iii) In addition, when needed, a source node S can set up a shared secret key K_{si} with router f_i using a well-studied key exchange protocol, e.g., Diffie-Hellman as in Passport [36]. This symmetric key exchange happens very infrequently thus representing only a one-time cost. Barak et al. [16] has proved that such a shared secret is *necessary* for any *secure* fault localization protocol via path monitoring.

3 $\mathrm{ShortMAC}$ Overview

We highlight the challenges of a secure fault localization protocol design, and then present our key ideas.

Challenge 1: Sophisticated packet modification attacks. In Fatih [43], WATCHERS [18, 25], and AudIt [10], each router records a *traffic summary* based on counters or Bloom Filters [17], which are updated with *no* secret keys for the packets the router forwards, and periodically exchanges local summaries with others for fault detection based on *flow reservation*. Without any authentication of the data packets, these schemes suffer from packet modification attacks. For example in AudIt [10], each router simply counts the number of packets it received for a certain path, and periodically sends the counter to the source node of the path for packet loss detection. However, malicious packet modification cannot be detected based only on the packet counts. Even when Bloom Filters are used [43] to reflect the packet contents, a malicious router can still tactically modify packets without affecting the Bloom Filter image (since Bloom Filters may not be collision-resistent).

Challenge 2: Colluding attacks. Routers in a path may employ "hop-by-hop" monitoring to detect packet delivery fault to reduce the communication overhead of sending the traffic summaries back to the source. For example in Figure 2, each router f_i asks for the traffic summaries (e.g., aknowledgements) *only from* the 2-hop neighbor f_{i+2} in the path, and accuses l_i if f_i does not receive the correct traffic summaries. In this approach however, if f_i is colluding with f_{i+1} and does not accuse f_{i+1} even if f_i does not receive the correct traffic summaries from f_{i+2}, then f_{i+1} can safely drop packets without being detected. Watchdog [41], Catch [40] and the proposal due to Liu et al. [34] are vulnerable to similar colluding attacks.

3.1 $\mathrm{ShortMAC}$ High-level Protocol Steps

To address the above challenges, $\mathrm{ShortMAC}$ monitors both the packet *count* and *content* at each hop. Specifically, a router maintains per-path counters to record the number of received data packets originated from the source in the current epoch. To ensure that the packet count is a verifiable measure of the desired monitoring task, we require that both packet modification and injection by malicious (colluding) routers affect counter values at benign nodes.

At the beginning of each epoch denoted by e_k, a source node S selects a path p and starts sending packets along p, with each packet carrying several $\mathrm{ShortMAC}$ authentication bits. The routers verify the authentication bits in each received packet based on the symmetric key shared with the source node, increment locally stored counters for p accordingly, and forward only the authentic packets. Due to the $\mathrm{ShortMAC}$ authentication bits, modified/injected packets can result in an observable deviation in the counter values which enable fault localization by the source at the end of each epoch.

At the end of each epoch e_k, the source S retrieves the counter reports from all routers and the destination in p for e_k, via a secure channel as Section 4 will describe. S then performs fault detection based on the retrieved counters, and bypasses the detected faulty link (if any) by finding another path excluding the identified faulty link (e.g., via source routing, path splicing [44], pathlet routing [21], or SCION routing [58]). The detection result is only used by S itself for selecting its own routing paths, instead of being shared with other nodes which is susceptible to framing attacks.

Although the high-level *epoch-based* protocol flow (nodes periodically send certain locally logged traffic summaries to the source) bears great similarity with Fatih [43], AudIt [10], and Statistical FL with sketch [16], both Fatih and AudIt use simple counters or Bloom Filters without *keyed* hash functions as the traffic summaries, thus remaining vulnerable to packet modification/injection attacks. In addition, the sketch-based packet fingerprints used in Statistical FL consume several hundreds of bytes *for each path*. In contrast, $\mathrm{ShortMAC}$ efficiently tackles packet modification attacks with only several-byte counters as shown below.

3.2 $\mathrm{ShortMAC}$ Packet Authentication

Our approach is to turn packet count into a reliable measure of packet content so that routers only need to store space-efficient counters. To this end, the integrity of the source's data packets must be ensured in order to detect malicious packet modification during the forwarding path; otherwise, a malicious router can always perform packet modification attacks without affecting the counter values, or inject bogus packets on behalf of the source to manipulate the counter values of the reporting routers. Hence, we reduce the problem to how the source node can authenticate its packets to all the routers in the path. However, traditional broadcast authentication schemes provide high authenticity for every *single* message, which is neither necessary nor practical in our setting where the messages are line-rate packets:

1) Not practical: On one hand, perfectly ensuring the authenticity of *every single* data packet introduces high overhead in a high-speed network. For example, digital signatures or one-time signatures for per-packet authentication is either computationally expensive or bandwidth-exhaustive, and using amortized signatures would either fail in the presence of packet loss or incur high communication overhead [38]. Attaching a Message Authentication Code (MAC) for each node along the path (as is used by Avramopoulos et al. [12]) is too bandwidth-expensive (e.g., reserving a 160-bit MAC space for each hop). In addition, TESLA authentication [48] would require time synchronization and routers to cache the received packets until the authentication key is later disclosed (longer than the end-to-end path latency). Finally, some recently proposed multicast/broadcast authentication schemes still require considerable communication overhead (e.g., up to hundreds of bytes per packet [39]) or multiple rounds for authenticating a message [19].

2) Not necessary: On the other hand, as we aim to *limit* the damage the adversary can inflict for a lower-bound guarantee on data-plane packet delivery, perfect per-packet authenticity is not necessary. Instead, our goal only requires the authenticity of a large fraction of data packets.

ShortMAC approach. Based on these observations, we propose ShortMAC, a light-weight scheme trading per-hop overhead with the adversary's ability to forge only a few (e.g., tens of) packets. More specifically, in ShortMAC, the source attaches to each packet a *k-bit random nonce, called k-bit MAC*, for each node on the path, where the parameter k is significantly less than the length of a typical MAC (e.g., $k = 2$). To construct the k-bit MAC for f_i, the source S uses a Pseudo-Random Function (PRF) which constructs a k-bit string as a function of the packet m and key K_{si} shared between S and f_i. We rely on the result that the output k-bit MAC is indistinguishable from a random k-bit string to any observer without the secret key K_{si} [42]. Each router f_i maintains two path-specific counters C_i^{good} and C_i^{bad} to record the numbers of received packets along that path with correct and incorrect k-bit MACs, respectively, in the current epoch. Such a scheme considerably reduces communication overhead compared to attaching entire MACs while retaining high security assurance and communication throughput, as shown later.

3.3 ShortMAC **Example**

We present a toy example in Figure 3 to provide intuition on how ShortMAC enables data-plane fault localization. Suppose the source node sends out 1000 packets in a certain epoch. The source uses a PRF taking a secret key as input which can map a packet into two bits (called 2-bit MAC) uniformly at random to anyone without knowledge of the secret key. The source computes the PRF four times for each packet, taking as input the epoch symmetric key shared with f_1, f_2, f_3, and the destination, respectively. Then the source attaches the resulting four 2-bit MACs to each packet.

Among the 1000 packets, suppose three packets are spontaneously dropped on the first link, and router f_1 receives the remaining 997 packets. f_1 computes the PRF on each of the received packets taking as input the epoch symmetric key shared with the source, and compares the resulting 2-bit MACs with the one embedded in each packet. All verifications are successful, so f_1 has $C_1^{good} = 997$ and $C_1^{bad} = 0$. Suppose the malicious router f_2 drops 100 good packets and injects 100 malicious packets. For each injected packet, f_2 needs to forge 2-bit MACs for both f_3 and the destination that "authenticate" the fabricated data content. However, since f_2 does not know the corresponding epoch symmetric keys of f_3 and the destination, f_2 can only guess the 2-bit MACs for its injected packets. Since the 2-bit MACs produced by the PRF are indistinguishable from random bits, f_2 can correctly guess each 2-bit MAC with probability $\frac{1}{4}$. Since f_2 must guess two correct MACs, each forged packet will be accepted by the destination with probability $\frac{1}{16}$. Suppose next that 26 of the 100 2-bit MACs that f_2 forged for f_3 happen to be valid with respect to the the malicious data content. f_3 thus computes $C_3^{bad} = 100 - 26 = 74$ and $C_3^{good} = 997 - 100$ (dropped legitimate packets) $+$ 26 (bogus but undetected packets) $= 923$. Similarly, we can analyze the counters for the destination in Figure 3, assuming 7 out of the 26 received bogus packets happen to be consistent with their 2-bit MACs at the destination.

3.4 Fault Localization and Guaranteed θ

At the end of each epoch, routers and the destination report their counter values to the source using a secure transmission approach (detailed in Section 4). The source can identify excessive packet drops between f_m and f_{m+1} if the C_{m+1}^{good} value of f_{m+1} is abnormally lower than that of f_m based on the drop detection threshold T_{dr} that is carefully set based on the customized acceptable per-link drop rate. Moreover, this scheme can successfully bound the total number of spurious packets with fabricated k-bit MACs

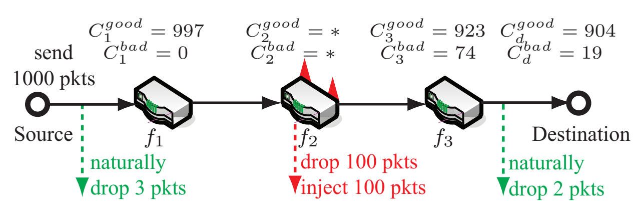


Figure 3. Fault localization example with ShortMAC **using 2-bit MAC. f_2 is malicious.**

that the adversary can inject, because at least one of the downstream recipient routers will detect the inconsistency of the k-bit MACs with a non-trivial probability, thus having a non-zero C^{bad} value. For example in Figure 3, although f_2 can claim any values for its own counters, no matter what values f_2 claims, the source can notice excessive packet loss and a large number of fake packets either between f_1 and f_2, or f_2 and f_3. Hence one of f_2's malicious links will be detected by the source.

Once the source S bypasses all malicious links identified by $\mathrm{ShortMAC}$, S can find a working path with no excessive packet corruption at any link, thus achieving a guaranteed successful forwarding rate θ. With secure fault localization, a source can find a working path after exploring at most Ω paths, where Ω is the number of malicious links in the network. In contrast, with only end-to-end path monitoring, a source may explore a number of paths exponential to Ω as we showed in Section 1.

4 $\mathrm{ShortMAC}$ Details

In this section we describe the $\mathrm{ShortMAC}$ protocol in detail, where the source can either guarantee that a high fraction θ of its data has been correctly forwarded if no malicious activities are detected, or can bypass the faulty links and find a working path after exploring a number of paths linear to the number of faulty links.[1] In the following, we first formalize the $\mathrm{ShortMAC}$ packet format and then detail the protocol.

4.1 $\mathrm{ShortMAC}$ Packet Format

A source node S adds a trailer to each data packet it sends:

$$trailer = \langle SN, \mathcal{M}_1, \ldots, \mathcal{M}_d \rangle, \quad (1)$$

where SN is a *per-path* sequence number to make each packet unique along the same path to prevent packet replay attacks, and $\mathcal{M}_i$ denotes the k-bit MAC computed for f_i, which is constructed in a recursive way starting from f_d:

$$\begin{aligned} \mathcal{M}_d &\leftarrow PRF_{K_{sd}}(IP_{invar}||SN||TTL_d) \\ \mathcal{M}_{d-1} &\leftarrow PRF_{K_{s(d-1)}}(IP_{invar}||SN||TTL_{d-1}||\mathcal{M}_d) \\ &\ldots\ldots \\ \mathcal{M}_i &\leftarrow PRF_{K_{si}}(IP_{invar}||SN||TTL_i||\mathcal{M}_{i+1}||\ldots||\mathcal{M}_d) \end{aligned} \quad (2)$$

where "$||$" denotes concatenation and $PRF_{K_{si}}(\cdot)$ denotes a PRF keyed by the symmetric key K_{si} shared between S and f_i. As previously discussed, the output of this PRF can be guessed correctly with probability no larger than $\frac{1}{2^k}$ by anyone without the secret key K_{si} [42]. In addition,

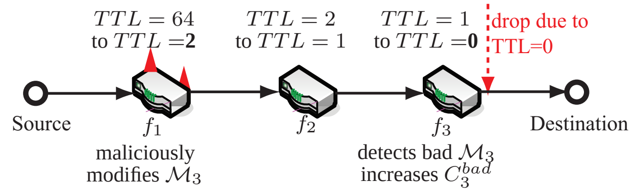


Figure 4. Illustration of framing attacks. f_1 is malicious.

1) IP_{invar} denotes the invariant portion of the original IP packet that should not be changed at each router during forwarding, including the packet payload and IP headers excluding variable fields such as `TTL`, `RecordRoute` IP option, `Timestamp` IP option etc. If these invariant fields are unexpectedly changed during forwarding, each downstream router can detect inconsistency between the (modified) packet and embedded k-bit MAC with a non-trivial probability $1 - \frac{1}{2^k}$ and thus increase its C^{bad} counter.

2) TTL_i denotes the expected TTL value at router i. Without authenticating this field in the k-bit MAC, a malicious router can strategically lower the TTL field to cause packet drop at a remote downstream router due to zero TTL value, thus performing framing attacks. For example in Figure 4, if $\mathcal{M}_i$ in Eq.(2) had not authenticated the TTL field, f_1 can maliciously change the TTL value in the packets to 2, instead of decrementing it by 1. This causes the packets to be dropped at f_3, thus framing the link between f_2 and f_3.

3) $\mathcal{M}_i$ also authenticates the downstream $\mathcal{M}_{i+1}, \ldots, \mathcal{M}_d$, so that if a malicious router f_m changes any of these downstream k-bit MACs, f_i can observe the inconsistency in $\mathcal{M}_i$ with a probability $1 - \frac{1}{2^k}$ and increase its C_i^{bad} value. Otherwise, the protocol is vulnerable to framing attacks. For example in Figure 4, if $\mathcal{M}_i$ in Eq.(2) had not authenticated the downstream $k-$bit MAC field, f_1 can maliciously modify $\mathcal{M}_3$ in the packets which causes f_3 to detect inconsistent $\mathcal{M}_3$ with a non-trivial probability and increase C_3^{bad}, thus framing the link between f_2 and f_3.

4.2 Protocol Details

Formally, $\mathrm{ShortMAC}$ consists of Request, Report, Identify, Bypass and Send stages, described as follows.

Stage 1: Request with hop-by-hop reliable transmission At the end of each epoch e_k (i.e., after sending every N data packets), the source S will send a request packet, denoted by `request`$= (S, p)$, along the path $p = (f_1, \ldots, f_d)$ used in epoch e_k. This `request` asks each router f_i and the destination f_d to report their counter values (C_i^{bad} and

[1] It has been proved that forwarding fault localization protocols protocols can only identify faulty links, rather than identifying the nodes [16]. However, given that a malicious node has a limited degree, after bypassing all its malicious links the source can eventually bypass that node.

C_i^{good}) along the reverse of path p. Then S expects these counter reports in Acknowledgment (ACK) packets from all the nodes in p containing the requested information authenticated with each node's K_{si}.

Note that a spontaneous loss of `request` or ACK packets will prevent S from learning the counter values by certain routers in the previous epoch. To preclude such damage, we use the following **hop-by-hop reliable transmission** approach: when f_i forwards either a `request` or an ACK packet to its neighbor, f_i tries up to r times (e.g., $r = 5$) until it gets a confirmation from the neighbor. In this way, the failure of receiving a `request` or ACK packet can only indicate malicious drops – more precisely, with the probability of $1 - \rho^r$, where ρ is the natural loss rate of a link. Then thanks to the Onion ACK approach presented below, the source can immediately identify a malicious link that drops or modifies `request` or ACK packets; hence the `request` packets do not need to be authenticated by the source as we show below.

Stage 2: Report with Onion ACK

Upon receiving a `request`, f_i starts a timer whose value is the maximum round trip time from f_i to the destination.[2] At the same time, f_i constructs its local report $\mathcal{R}_i$:

$$\mathcal{R}_i = \left(f_i, p, C_i^{good}, C_i^{bad}\right) \tag{3}$$

where f_i is the node id, p is the requested path, and C_i^{good} and C_i^{bad} are the counter values from the previous epoch. Each router finds C_i^{good} and C_i^{bad} corresponding to path p based on the source and destination IDs in p (assuming single path routing). Once the report is constructed:

Case 1 If f_i receives an ACK $\mathcal{A}_{i+1}$ from neighbor f_{i+1} before the timer expires, f_i further commits $\mathcal{R}_i$ into a new ACK $\mathcal{A}_i$ by combining the received $\mathcal{A}_{i+1}$ via an **Onion ACK** approach:

$$\mathcal{A}_i = \left(\mathcal{R}_i, \mathcal{A}_{i+1}, \mathrm{H}_{K_{si}}(\mathcal{R}_i || \mathcal{A}_{i+1})\right), \tag{4}$$

$\mathrm{H}_{K_{si}}(\cdot)$ denotes a message authentication code computed with K_{si}. Then, f_i forwards $\mathcal{A}_i$ to f_{i-1} toward S.

Case 2 If f_i receives no ACK packet from f_{i+1} before the timer expires, f_i will initiate a new ACK with its local report and send it to f_{i-1}.

The Onion ACK prevents the adversary from *selectively* dropping the `request` or the reports of a certain router f_i and framing a benign link l_i [59]. In Onion ACK, all the reports are *combined* and *authenticated* in one ACK packet at each hop so that a malicious node can only drop or modify the onion report from its *immediate* neighbors. Intuitively, if f_m drops or modifies the received `request` or Onion ACK, the source can receive the correct reports from $f_1, \ldots, f_{m-1}$ but not from $f_m, \ldots, f_d$; hence one of f_m's links will be pinpointed by the source node, in the identify stage described below.

After sending the local reports, each router f_i resets C_i^{good} and C_i^{bad} to zero, to be used for the next epoch along path p (if p is still used).

Stage 3: Identify

Upon receiving an Onion ACK $\mathcal{A}_1$ from f_1, S first iteratively retrieves $\mathcal{A}_1, \mathcal{A}_2, \ldots$ in order, until it either completes at d or fails at j ($j \neq d$).[3] When the check fails at j ($j \neq d$), S will immediately identify l_j as faulty due to the use of reliable hop-by-hop transmission and Onion ACK. For example, if S receives no report it will identify l_1 as faulty ($j = 1$).

In addition, S extracts $\mathcal{R}_1, \ldots, \mathcal{R}_j$ in turn which include the C_i^{bad} and C_i^{good} values. A non-zero C_i^{bad} implies the existence of malicious packet injection between f_i and S. However, S cannot blame l_i simply whenever $C_i^{bad} > 0$, say, $C_i^{bad} = 1$. A possible scenario is that a malicious node f_{i-2} injects a fake packet, but the k-bit MAC intended for f_{i-1} "happens" to be consistent with the fake packet at benign node f_{i-1} (e.g., when $k = 2$, this can happen with probability 0.25). In this case, f_{i-1} will forward the fake packet which f_i may detect and thus increase C_i^{bad}. Similarly, due to natural packet loss, S cannot simply accuse link l_i when $C_i^{good} < C_{i-1}^{good}$. Therefore, we leverage two detection thresholds T_{in} and T_{dr}, where T_{in} is the injection detection threshold for the number of injected packets on each link, and T_{dr} is the drop detection threshold for the fraction of dropped packets on each link. As we will show in Section 6, these thresholds reduce false positives while limiting the adversary's ability to corrupt packets and ensuring a lower bound on the successful packet forwarding rate. The detection thresholds are used in two detection procedures:

1) $\mathrm{check\text{-}injection}$**:** S checks the extracted $C_1^{bad}, C_2^{bad}, \ldots$, C_j^{bad} values *in order*. If $C_i^{bad} \geq T_{in}$ for some i, then S identifies l_i as faulty and the $\mathrm{check\text{-}injection}$ procedure stops.

2) $\mathrm{check\text{-}dropping}$**:** If no fault is detected by $\mathrm{check\text{-}injection}$, S further checks the extracted $C_1^{good}, C_2^{good}, \ldots$, C_j^{good} values *in order*. If $C_i^{good} < (1 - T_{dr}) \cdot C_{i-1}^{good}$ (with $C_0^{good} = N$) holds for certain i, then S identifies l_i as faulty and the $\mathrm{check\text{-}dropping}$ procedure terminates.

Stage 4: Bypass and Send

If Stage 2 outputs any malicious link l_m, S selects a new path excluding the previously detected malicious links and sends its packets with $\mathrm{ShortMAC}$ authentication shown

[2] We can expect a reasonable upper bound of link latency in benign cases, which can be used to compute the maximum round trip time according to the hop count from f_i to the destination. Avramopolous et al. [12] first introduced the use of such a timer.

[3] S can verify if a certain retrieved report $\mathcal{R}_i$ is valid by checking the embedded message integrity code $\mathrm{H}_{K_{si}}(\mathcal{R}_i || \mathcal{A}_{i+1})$.

in Eq.(2). Each node f_i examines its corresponding k-bit MAC $\mathcal{M}_i$ in each packet to increase C_i^{good} or C_i^{bad} accordingly. In addition, each router remembers the last seen *per-path* SN embedded in the packets as shown in Eq.(1), and discards packets with older SN in that path.

5 Security Analysis

This section discusses $\mathrm{ShortMAC}$'s security against data-plane attacks by malicious routers. Section 6 provides theoretical proofs on $\mathrm{ShortMAC}$'s security. In our adversary model, a malicious router can drop and inject data packets, `request`s and `ACK`s, and can send arbitrary counter values in its reports. We show that $\mathrm{ShortMAC}$ is secure against a single malicious router (say, f_m) as well as multiple colluding nodes.

Corrupting data packets. Dropping legitimate data packets by f_m will cause a discrepancy of the counter values between f_m and its neighbors. For example, if f_m correctly reports C_m^{good}, then $C_m^{good} - C_{m+1}^{good}$ will exhibit a large discrepancy; if f_m reports a lower C_m^{good}, then $C_{m-1}^{good} - C_m^{good}$ will exhibit a large discrepancy. Hence, either l_{m-1} or l_m will become suspicious. Moreover, if f_m injects/modifies packets, $\mathcal{M}_{m+1}$ will be inconsistent at f_{m+1} with high probability and cause a non-zero C_{m+1}^{bad}. Hence, both dropping and injection attacks can be detected as long as the source can learn the correct counter values in the `ACK` packets sent by the nodes between f_m and the destination, which is described next.

Corrupting `ACK`s or `request`s. Since the `request`s are not authenticated by S, f_m can modify the content of `request`s (such as the source ID and the path); however, this will result in S failing to receive the correct counter reports from f_{m+1} (or f_m) $, \ldots, f_d$ in p, thus causing l_{m+1} or l_m to be detected. f_m cannot selectively drop the `ACK` reports due to the use of Onion `ACK`. Instead, f_m can only drop the `ACK`s or `request`s from its *immediate* neighbors, which will again harm its incident links.

Replay, reorder, and traffic analysis attacks. To prevent replay and reorder attacks, each packet contains a per-path sequence number SN in Eq.(1) and each router discards packets with older SNs. Hence, the replayed and reordered packets will be dropped at the next-hop benign node without influencing the counter values of benign nodes. Note that because $\mathrm{ShortMAC}$ runs on a per-path basis and a SN is a *per-path* sequence number providing natural isolation across different paths, packets *along the same path* are expected to maintain the same order during forwarding as they were sent by the source in benign cases. On the other hand, if f_m falsely reports a large SN, f_{m+1} will drop the subsequent packets and l_m will be identified as malicious due to its high packet drop rate. Moreover, the per-path SN can prevent $\mathrm{ShortMAC}$ from *traffic analysis attacks*, where f_m attempts to find out the correct k-bit MAC of a packet m by re-sending m with different k-bit MACs and observing whether the next-hop f_{m+1} forwards the packet. Such traffic analysis is ineffective because f_{m+1} can detect packets with the same SN and each packet is unique due to the use of the per-path SN, and thus f_m cannot send the same packet m with only the k-bit MAC changed.

DoS attacks. A malicious router f_m may launch bandwidth Denial-of-Service (DoS) attacks by generating an excessive amount of packets. However, this attack can be reduced to a packet injection attack and will be reflected by C_{m+1}^{bad}. A malicious router may also attempt to open many bogus flows with spoofed sources to exhaust other routers' state. We can borrow existing work to provide source accountability and reliable flow/path identification [8, 56]. Also note that in our adversary model we consider malicious *routers* which threaten the communication between benign hosts. We do not consider DDoS attacks launched by malicious hosts (botnets), which other researchers have strived to defend against [35, 37, 56]. Hence in our problem setting, a link under DDoS attacks thus exhibiting high loss rate is simply considered a faulty link under our adversary model. Meanwhile, the path setup phase in $\mathrm{ShortMAC}$ can be naturally integrated with capability schemes [56] for DDoS limiting, and the per-path counters may also be used for per-path rate limiting.

Collusion attacks. Each of the colluding routers can commit any of the misbehavior discussed above. We can prove by induction that in any case, one of the malicious links of one of the colluding nodes is guaranteed to be detected. A proof sketch is given below.

Consider the base case where two nodes f_m and $f_{m'}$ ($m < m'$) collude. Without loss of generality:

1) When f_m and $f_{m'}$ are not adjacent (i.e., $m' > m + 1$), the security analysis in Section 5 applies to f_m and one of f_m's malicious links will become suspicious if f_m misbehaves. This is because if f_m commits the above attacks, such misbehavior will be reflected in the benign neighbor f_{m+1}'s counters which cannot be biased by $f_{m'}$.

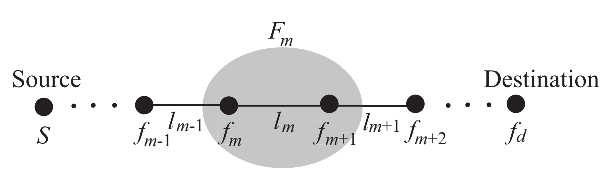


Figure 5. Security against colluding nodes – one base case with two adjacent colluding nodes f_m and f_{m+1} forming a virtual malicious node F_m.

2) When f_m and $f_{m'}$ are adjacent ($m' = m + 1$), these two nodes can be regarded as one single "virtual" malicious node F_m with neighbors f_{m-1} and f_{m+2}, as shown in Figure 5. (i) If f_m or f_{m+1} drops packets, a discrepancy will exist between C^{good}_{m-1} and C^{good}_{m+2}, no matter what values of C^{good}_{m} and C^{good}_{m+1} F_m claims. (ii) If f_m or f_{m+1} injects packets, C^{bad}_{m+2} will become non-zero and make l_{m+1} suspicious. In any case, an adjacent link of F_m (a malicious link) will become suspicious.

In the general case with n colluding nodes, we can first group adjacent colluding nodes into virtual malicious nodes as in Figure 5, resulting in non-adjacent malicious nodes (including virtual malicious nodes). Then we can show non-adjacent malicious nodes can be detected based on the above analysis.

Despite colluding attackers cannot corrupt packets more than the same thresholds as an individual attacker on any *single* link, they can choose to distribute packet dropping across multiple links. In this case, the total packet drop rate by colluding attackers increases (and is still bounded) linearly to the number of malicious links in the same path, as analyzed in Section 6.

6 Theoretical Results and Comparison

We prove the $(N, \delta)-$data-plane fault localization (Definition 4) and $(\alpha, \beta)_\delta-$forwarding security of $\mathrm{ShortMAC}$ (Definition 6), which in turn yield the $\theta-$guaranteed forwarding correctness (Definition 5). Proofs of the lemmas and theorems are provided in Appendix A.

Comparison of theoretical results. Before presenting the theorems, we first summarize and compare $\mathrm{ShortMAC}$ theoretical results with two recent proposals, PAAI-1 [59] and Stat. FL [16] (including two approaches denoted by *SSS* and *sketch*). Table 1 presents the numeric figures using an example parameter setting for intuitive illustration, while $\mathrm{ShortMAC}$ presents similarly distinct advantages in other parameter settings. In this example scenario shown in the table, the guaranteed data-plane packet delivery ratio is $\theta = 92\%$. The communication overhead for a router in $\mathrm{ShortMAC}$ is 1 extra ACK for every 3.8×10^4 data packets in an epoch; the marking cost is 10 bits for the 2-bit MACs in a path with 5 hops, and the per-path state at each router is 21 bytes (16-byte symmetric key, 2-byte C^{good}, 1-byte C^{bad}, and 2-byte per-path SN). Though Barak et al. *proved the necessity* of per-path state for a *secure* fault localization protocol [16], such a *minimal* per-path state in $\mathrm{ShortMAC}$ is viable for both intra-domain networks with tens of thousands of routers and the Internet AS-level routing among currently tens of thousands of ASes.

We provide the intuition for $\mathrm{ShortMAC}$'s distinct advantages. PAAI-1 or Stat. FL used either low-rate packet sampling or approximation techniques for packet fingerprinting, both of which waste entropy contained in certain packet transmissions, thus resulting in long detection delay (e.g., the transmission results of non-sampled packets will not contribute to the detection phase). In contrast, $\mathrm{ShortMAC}$ counts *every* packet transmission thus achieving much faster detection rate. In addition, *secure* packet sampling requires additional packet buffering [59], and packet fingerprint takes considerable memory [16].

Lemma 1 ***Injection Detection:*** *Given the bound δ on detection false negative and false positive rates, the injection detection threshold T_{in} can be set to $T_{in} = \frac{2\ln\frac{2d}{\delta}}{q^4}$, where d is the path length and $q = \frac{2^k-1}{2^k}$ is the probability that a fake packet will be inconsistent with the associated k-bit MAC. The number of fake packets β an adversary can inject on one of its malicious links without being detected is limited to: $\beta = \frac{T_{in}}{q} + \frac{\sqrt{\left(\ln\frac{2}{\delta}\right)^2 + 8qT_{in}\ln\frac{2}{\delta}} + \ln\frac{2}{\delta}}{4q^2}$.*

In Lemma 2, we derive N, the number of data packets a source needs to send in one epoch to bound the detection false positive and false negative rates below δ. Due to natural packet loss, a network operator first sets an expectation based on her domain knowledge such that any benign link in normal condition should spontaneously drop less than ρ fraction of packets. We first describe how the drop detection threshold T_{dr} is set when N and δ are given. Intuitively, by sending more data packets (larger N), the *observed* per-link drop rate can approach more closely its *expected* value, which is less than ρ; otherwise, with a smaller N, the observed per-link drop rate can deviate further away from ρ, and the drop detection threshold T_{dr} has to tolerate a larger deviation (thus being very loose) in order to limit the false positive rate below the given δ. On the other hand, a small N is desired for *fast fault localization*. We define **Detection Delay** to be the minimum value of N given the required δ.

Lemma 2 ***Dropping Detection*** *and (N, δ)-* ***Fault Localization:*** *Given the bound δ on detection false positive and negative rates and drop detection threshold T_{dr}, the detection delay N is given by: $N = \frac{\ln(\frac{2d}{\delta})}{2\left(T_{dr}-\rho\right)^2\left(1-T_{dr}\right)^d}$, where d is the path length. Correspondingly, the fraction of packets α an adversary can drop on one of its malicious links without being detected is limited to: $\alpha = 1 - (1 - T_{dr})^2 + \frac{\beta}{N(1-T_{dr})^d}$.*

In practice, T_{dr} can be chosen according to the expected upper bound ρ of a "reasonable" normal link loss rate such that a drop rate above T_{dr} is regarded as "excessively lossy".

Theorem 1 ***Forwarding Security and Correctness:*** *Given T_{dr}, δ, and path length d, we can achieve*

Protocol	ShortMAC	PAAI-1	SSS	Sketch
Detect. Delay (pkt)	3.8×10^4	7.1×10^5	1.6×10^8	$\approx 10^6$
Comm. (extra %)	$< 10^{-5}$	1	1	$< 10^{-5}$
Marking Cost (bytes)	2	0	0	0
Per-path State (bytes)	21	$2{\times}10^5$	4×10^3	≈ 500

Table 1. Theoretical comparison with PAAI-1 [59] and Stat. FL [16] (including two approaches *SSS* and *sketch*). Note that the details of *sketch* are not provided in the published paper [16], and the full version of [16] does not present the explicit bounds on detection delay. The above figures for *sketch* are estimated from their earlier work [?]. In this example scenario, d = 5, $\delta = 1\%$, $\rho = 0.5\%$, $T_{dr} = 1.5\%$, a symmetric key is 16 bytes, and ShortMAC uses 2-bit MACs. PAAI-1 specific parameters include the "packet sampling rate" set to 0.01, the end-to-end latency set to 25 ms, the source's sending rate set to 10^6 packets per second, each packet hash is 128 bits.

$(\alpha,\beta)_\delta-$*forwarding security where* α *is given by Lemma 2 and* β *is given by Lemma 1. We also achieve* (Ω,θ)*-Guaranteed forwarding correctness with* Ω *equal to the number of malicious links in the network, and* $\theta = (1-T_{dr})^d - \frac{\beta}{N}$*. where* N *is derived from Lemma 2*

In Theorem 2, we analyze the protocol overhead with the following three metrics (we further analyze the *throughput* and *latency* in Section 8 via real-field testing):

1) The ***communication overhead*** is the fraction of extra packets each router needs to transmit.

2) The ***marking cost*** is the number of extra bits a source needs to embed into each data packet.

3) The ***per-path state*** is defined as the per-path extra bits that a router stores for the security protocol in *fast memory* needed for *per-packet* processing.[4]

Theorem 2 ***Overhead:*** *For each router, the communication overhead is one packet for each epoch of* N *data packets. The marking cost is* $k \cdot d$ *bits for the* k*-bit MACs where* d *is the path length. The per-path state comprises one* $\lg N$*-bit* C^{good} *counter, one* $\lg \beta$*-bit* C^{bad} *counter, one* $\lg N$*-bit last-seen per-path* SN*, and one epoch symmetric key.*

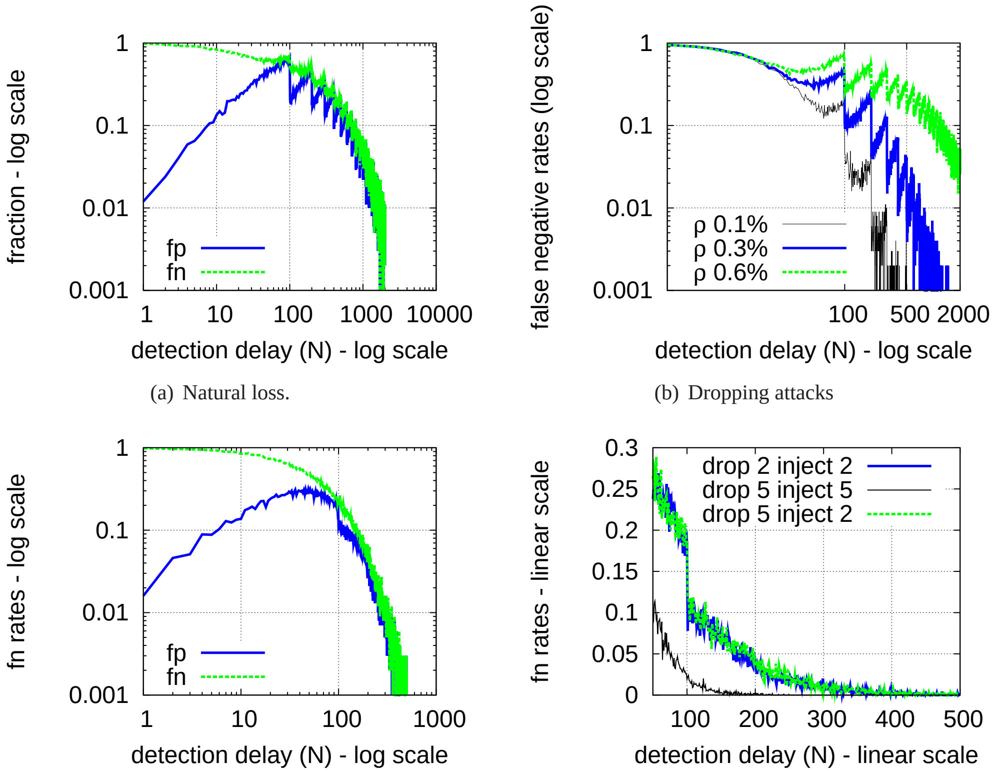
7 SSFNet-based Evaluation

In addition to analyzing the theoretical performance, we implement ShortMAC prototype on the SSFNet simulator [6] to study the detection delay and security of ShortMAC. Section 8 further investigates ShortMAC's throughput and latency. These experimental results provide *average-case* performance with various attack strategies to complement the theoretical results derived in the *worst case* scenario (due to multiple mathematical relaxations such as Hoeffding inequality) and constant dropping/injection rates.

Evaluation scenario and attack pattern. Since ShortMAC provides a natural isolation across paths due to its per-path state, our evaluation focuses on a single path. Specifically, we present the result of a 6-hop path (routers f_1, f_2, f_3, f_4, f_5 and the destination f_6) since our experiment yields the same observation with other path lengths. We simulate both an (i) *independent packet corruption* pattern where a malicious node drops/injects each packet independently with a certain drop/injection rate, and (ii) *random-period packet corruption* pattern where the benign (non-attack) period T_b and attack period T_a (when the malicious node drops/modifies *all* legitimate packets) are activated in turns. The durations for both periods are randomly generated. For both attack patterns, we control the *average* packet drop/injection rates and observe that both attack patterns yield similar observations. Hence, in the following experiment, we only show the results for the independent packet corruption pattern. Also, we infuse natural packet loss rate ρ for each link to simulate natural packet loss, which is not provided by SSFNet. As Section 5 elaborates ShortMAC security against colluding attacks, we only show the representative results for a single malicious node f_3. For each simulation setting, we run the simulation 1000 times and present the average results.

Against various dropping attacks. Figure 6(a) depicts the detection delay N and error rates δ with per-link natural loss rate ρ as 0.5%, drop detection threshold T_{dr} as 1%, and a stealthy malicious drop rate as 2%. We see that even against stealthy dropping attacks with a dropping rate as low as 2%, ShortMAC can successfully localize a faulty link in < 2000 packets with an error rate $\delta < 1\%$, which is orders of magnitudes faster than the worst-case theoretical bound (Lemma2). Figure 6(b) depicts different detection delays

[4]The buffering space needed for the Onion-ACK construction of `report` messages in ShortMAC is not a major concern, as the Onion-ACK is computed only once every epoch, which can be buffered in off-chip storage.



(a) Natural loss.

(b) Dropping attacks

(c) Injection attacks.

(d) Combined attacks.

Figure 6. In this set of simulations, f_3 is the malicious router performing attacks. The parameter are set as follows: (a) The malicious drop rate is 2%, $T_{dr} = 1\%$, and natural drop rate $\rho = 0.5\%$. (b) The malicious drop rate is 2%, and $T_{dr} = 1\%$. (c) The malicious injection rate is 2% using 2-bit MACs, natural loss rate $\rho = 0.5\%$, and $T_{dr} = 1\%$. (d) "drop p inject q" denotes the use of $p\%$ dropping rate and $q\%$ injection rate at f_3.

with different natural packet loss rates, demonstrating that larger $|T_{dr} - \rho|$ yields higher detection accuracy and lower detection delay.

Against various injection attacks. Figure 6(c) shows the results when f_3 injects packets at a 2% rate (relative to the legitimate packet sending rate). It shows that the error rates stay below 1% in a few hundred packets, indicating that even with 2-bit MACs, an adversary can only inject up to around ten packets without being detected. We further investigate the effects of using different lengths of $k-$bit MACs, and Figure 7 shows that the detection delay and error rate dramatically diminish as k increases.

Against combined attacks. Figure 6(d) shows how the combinations of dropping and injection attack strategies (in our setting, dropping/injection rates are chosen between 2% – 5%) influence the protocol. We observe that the detection delay is mainly determined by the dropping detection process, which is much slower than the injection detection process. This also indicates that a malicious node cannot gain any advantage (and actually can only harm itself) by injecting bogus packets in attempt to bias the counter values.

Variance due to different malicious node positions. To investigate the influence of the position of the malicious node, we consider a path with 6 forwarding nodes $f_1, f_2, \ldots, f_6$ and place the malicious node at each position (1 to 6) in turn. We limit the error rate $< 1\%$ and obtain the corresponding detection delays. Figure 8 shows one representative scenario where both dropping and injection rates are 5%. We can see that (i) the dropping detection delay increases linearly when the malicious node is farther away from the source. This is because in the $\mathrm{ShortMAC}$ detection process, the source always inspects the closer links first and stops once the first "faulty" link is detected. The FP rate thus increases when more links exist between the source and the malicious node due to natural packet loss on each link. (ii) In contrast, the injection detection delay exhibits little variance (cannot be seen from the figure as the detection delay is determined by the dropping detection), which can

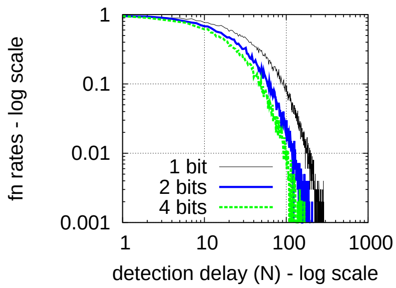


Figure 7. Effects of different k-bit MAC lengths on detection delay N and false negative rate δ. The malicious injection rate is 2%, $\rho = 0.5\%$, and $T_{dr} = 1\%$.

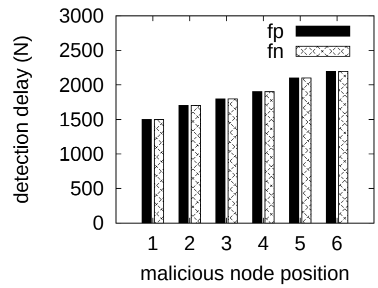


Figure 8. Variance on detection delay N in dropping attacks. $\delta < 1\%$, $T_{dr} = 1\%$, $\rho = 0.5\%$, and both malicious dropping and injection rates set to 5%.

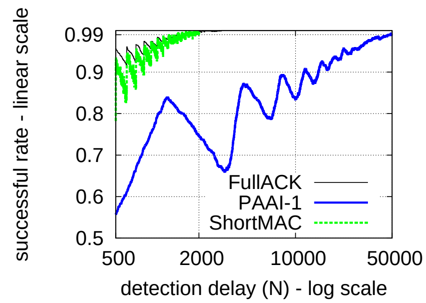


Figure 9. Comparison with PAAI-1 and FullACK [59]. The natural packet loss rate $\rho = 0.5\%$ and drop detection threshold $T_{dr} = 1\%$.

	ShortMAC	FullACK	PAAI-1
Detect. delay	20 sec	20 sec	8.3 min
Communication	0.01%	100%	5.6%

Table 2. Comparison of ShortMAC, FullACK, and PAAI-1 with a source send rate of 100 packets per second.

also be theoretically proved.

Comparison with recently proposed protocols. For comparison, we simulate the recently proposed FullACK and PAAI-1 [59] schemes presenting the lowest detection delays to date. FullACK is a heavy-weight fault localization protocol requiring an Onion ACK packet from *every* forwarding node for *every* packet the source sent. In contrast, PAAI-1 employs packet sampling and only requires acknowledgments for the securely sampled packets to reduce communication overhead while retaining desired detection delay. Since both FullACK and PAAI-1 only consider packet dropping attacks, we compare their dropping detection delays along a path with 6 hops and f_3 as the malicious node. Figure 9 shows the results when per-link natural packet loss rate $\rho = 0.5\%$ and drop detection threshold $T_{dr} = 1\%$. To make the comparison clear, we use a metric of *successful rate*, which equals to 1 - $max\{$FP rate, FN rate$\}$. The results show that the detection delays to achieve a successful rate $> 99\%$ for ShortMAC, FullACK, and PAAI-1 are 2000, 2000, and 5×10^4, respectively. Table 2 shows their detection delays in seconds/minutes and compares the extra communication overhead, based on the results from Figure 9 and with $\delta < 1\%$.

8 Linux Prototype and Evaluation

We implement ShortMAC source and destination nodes as user-space processes running on Ubuntu 10.04 32-bit Desktop OS. Even implemented in user-space on a standard desktop OS, our result shows that the cryptographic operations of ShortMAC incur little communication degradation and negligible additional latency at gigabit line rate. It has also been demonstrated that using modern hardware implementation and acceleration the speed of PRF functions can be fundamentally improved [29].

Implementation details. Our ShortMAC processes listen to application packets via TUN/TAP virtual interfaces and appending k-bit MACs to the packets. We also implement ShortMAC routers using the Click Modular Router [28] running on Ubuntu 10.04 32-bit Desktop OS, which verify the k-bit MACs in each packet at each hop. To approach the realistic performance of commercial-grade routers, we implement the above elements on off-the-shelf servers with an Intel Xeon E5640 CPU (four 2.66 GHz cores with 5.86 GT/s QuickPath Interconnect, 256KB L1 cache, 1MB L2 cache, 12MB L3 cache, and 25.6 GB/s memory bandwidth) and 12G DDR3 RAM. The servers are equipped with Broadcom NetXtreme II BCM5709 Gigabit Ethernet Interface Cards.

Evaluation methodology. We evaluate ShortMAC's effects on communication throughput and computational overhead, especially due to the generation and verification of k-bit MAC using PRF operations. We utilize the widely used Netperf benchmark [4] for the ShortMAC throughput evaluation, and write our own micro-benchmark for accurate latency evaluation. We evaluate ShortMAC with varying packet sizes by configuring the interface Maximum Transmission Unit (MTU) sizes. We evaluate the through-

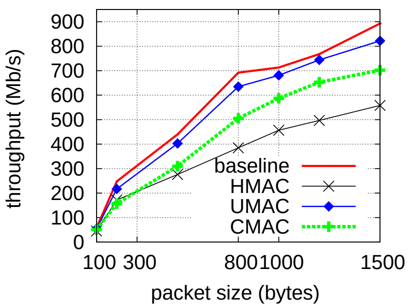


Figure 10. Router throughput.

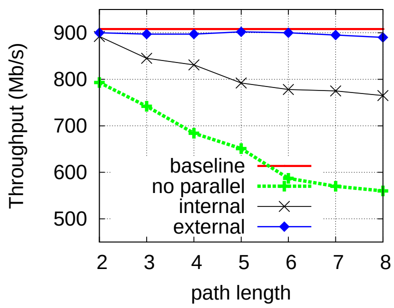


Figure 11. Source throughput.

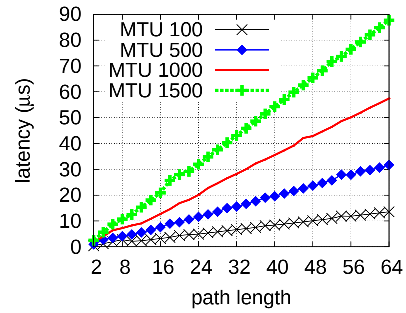


Figure 12. Source latency.

put of a ShortMAC router and a ShortMAC source separately to better illustrate the throughput of each component, while the end-to-end path throughput can be easily derived by taking the minimum throughput of the two evaluation results. Then we evaluate the end-to-end latency with different path lengths ranging from 2 to 64. We also exploit the multi-core parallel processing at the source node via OpenMP API [5].

Summary of evaluation results. The evaluation results of our Linux software prototype demonstrate that both a ShortMAC router and source node can retain more than 92% of the *baseline throughput* (no ShortMAC operations are employed). Furthermore, the additional latency due to ShortMAC operations is negligible (tens of microseconds) even with a path length of 64 hops. The results further indicate the ShortMAC scheme is fully scalable as the number of processing cores increases in a software-based implementation, while we anticipate hardware implementation of the MAC operations in ShortMAC can further boost the protocol throughput. Details of the evaluation results are as follows.

Router throughput with different PRF implementations. We first evaluate the throughput of a user-level ShortMAC router with different PRF implementations (i.e.,UMAC [51], HMAC-SHA1 [30], and AES-CMAC [50]) with the support of the new Intel AES-NI instructions [26]. The ShortMAC router connects a source machine and a destination machine, with the source sending TCP packets via Netperf as fast as possible to the destination to stress-test the router. For comparison, we use the Linux kernel forwarding throughput *without* ShortMAC operations as the base line. The ShortMAC router runs as a *single user-space* process without exploring parallelism, which already matches up the base line speed as shown below.

Figure 10 depicts the results with packet sizes from 100 to 1500 bytes, showing that UMAC-based PRF implementation yields the highest throughput, which retains more than 90% of the baseline throughput (e.g., 92% with 1.5KB packet size and 96% with 1KB packet size). With a small packet size of 100 bytes, both the baseline and ShortMAC throughput dropped substantially (similar to other public testing results [3]), because the network drivers used in our experiments are running under interrupt-driven mode, which hampers throughput when packet receiving rate is high. However, UMAC-based PRF still retains $\frac{53.84}{57.52}$=94% of the baseline throughput.

Source node throughput. We further evaluate the throughput of a ShortMAC source node with different path length d, where for each path length the source needs to perform $d-1$ UMAC-based PRF operations. Originally, it might seem that the ShortMAC source node represents the throughput bottleneck as the source needs to compute multiple k-bit MACs. However by parallelizing the ShortMAC operations on readily-available multi-processor systems, the throughput of a ShortMAC source node can *fully* cope with the base line rate even with a path length of 8. For comparison, we use the source node throughput *without* ShortMAC operations as the baseline. We evaluate two different parallelizations based on widely used OpenMP [5] API. Our first implementation (internal parallelism in short) uses multiple OpenMP threads to parallelize the computation of multiple k-bit MACs per packet. Our second implementation (external parallelism in short) assigns different packets to different OpenMP threads.

We evaluate the ShortMAC source throughput with various packet sizes, and observe that in all cases ShortMAC incurs negligible throughput degradation. Hence we only show the results with packet size set to 1500 bytes in Figure 11. We can see that external parallelism yields the best performance, which matches the baseline case where the source performs no ShortMAC operations.

ShortMAC latency. We also evaluate the additional latency incurred by a ShortMAC source node for computing the k-bit MICs with different path lengths and packet sizes; while the end-to-end latency can be derived base on our re-

Path Length	Checksum (μs)	UMAC (μs)			
		100	500	1000	1500
2	0.0374	0.1771	0.4760	0.8892	1.4047
3	0.0378	0.3691	0.9557	1.7635	3.3025
4	0.0442	0.5239	1.4273	2.6357	4.0944
5	0.0415	0.7080	1.9018	3.5059	5.4566
6	0.0437	0.8723	2.3758	4.3839	6.8307
7	0.0445	1.0467	2.8530	5.2617	8.2019
8	0.0474	1.2206	3.3274	6.1285	9.5483

Table 3. ShortMAC **source node latency breakdown (checksum updates and UMAC computation). All the data represent the average time of processing 50000 packets.**

sults. This additional latency in ShortMAC includes PRF computation, k-bit MICs appending, and TCP/IP checksum updating. We write our micro-benchmark to derive the additional time delay for the source to send each packet compared to the baseline case where the source does not compute any k-bit MIC nor updates the checksums.

Figure 12 and Table 3 show the results. We can see that the latency incurred by the checksum computation is stable. It does not increase with the packet size because in our implementation we employ incremental checksum update for the short MIC appended to the packet, instead of recomputing the checksum over the entire packet. We do not observe sharp increase of checksum latency with increasing path length either due to ShortMAC's efficient k-bit MIC authentication. In addition, the latency caused by the checksum computation is small compared to the latency introduced by UMAC-based PRF computation. The additional latency due to UMAC computation increases linearly to the path length under the same packet size, and also increases linearly to the packet size with a fixed path length due to the property of the UMAC algorithm. Finally, compared to the average end-to-end network latency which is on the order of milliseconds, the additional latency introduced by ShortMAC is negligible.

9 Discussion and Limitations

Incremental deployment. Although we argue it is feasible to upgrade all routers with ShortMAC within ISP/enterprise networks, we observe that partial deployment of ShortMAC can still provide benefits and thus enables incremental deployment. Specifically, the ShortMAC routers form an *overlay network* on top of the physical network. In the overlay network, a "logical link" consists of the physical links between two ShortMAC routers. The fault localization protocol runs only on the ShortMAC routers and a data delivery fault will be localized to a logical link. Although in such settings the source node cannot exactly identify a faulty physical link, it can nevertheless localize the fault to a network area (a set of links between two ShortMAC routers) to facilitate further investigation. Furthermore, the more densely the ShortMAC routers are deployed, the more accurate the fault localization can be, which incentivizes incrementally deploying ShortMAC. However, one caveat for incremental deployment is that a discovery protocol for determining which routers support ShortMAC is needed, possibly through the use of explorer packets.

Interdomain deployment. Though ShortMAC mainly targets at intra-domain networks such as ISP and enterprise networks, ShortMAC may also be deployed in interdomain networks such as the Internet. In the interdomain setting, each Autonomous System (AS) can represent a node in ShortMAC; the fault localization runs at the AS level and localizes any data delivery fault between two ASes. To make ShortMAC applicable, different ASes need to establish secret keys (e.g., via Passport [36]), and the egress router of an AS needs to set the TTL value of each packet to the TTL value at the ingress router *minus one* to enable k-bit MAC verification (Section 4.1). Finally, a source AS needs to know the downstream AS path (which is readily available in BGP) which may dynamically change in the current Internet; however, the majority of AS paths are stable over minutes [49] thus facilitating ShortMAC fault localization. If an adversary were to constantly alter paths, it would essentially raise suspicion to itself, since path information is visible and the adversary needs to remain on the path to remain effective.

Topology changes and short-lived flows. Fault localization protocols inevitably require at least a threshold number of packets to be sent along the monitored path to obtain a statistically accurate detection in the presence of natural packet loss. Hence, monitored paths need to be stable over an epoch. Since ShortMAC incurs several orders of magnitude lower detection delay compared to related work [16, 59], ShortMAC can support topology or path changes and short-lived flows much better than previous work. For example, as long as the path remains stable for transmitting around 2000 packets, the source can make an accurate fault localization. While path changes do happen during an epoch (e.g., due to link failures), the source will detect the old link where the path is switched away as faulty. At the same time, the source can also learn the routing updates about the path change, and by correlating the detection results with routing updates, the source may distinguish a benign path change and a malicious packet misrouting attack (in which case no corresponding routing updates will be received). However, the fault localization accuracy of ShortMAC decreases for dynamic paths that transmit far fewer than 2000 packets before path changes occur.

10 Related Work

Perlman described the idea of acknowledgment-based approaches to detect data-plane adversaries and achieve robust routing in the presence of Byzantine failures [47]. In Sprout [20], a source node monitors the end-to-end path performance and uses probabilistic route selection to find a working path if the current path is faulty. However, without secure fault localization, both schemes suffer from the exponential path exploration problem as Figure 1 shows. Given the importance of fault localization, several approaches have been proposed, which unfortunately suffer from the following limitations.

Security vulnerabilities. In ODSBR [14, 15] and Secure Traceroute [46], the source node monitors the end-to-end loss rate of the path; and only when the observed loss rate exceeds a certain threshold, the source starts probing specific nodes in the path soliciting acknowledgments for the *subsequent* packets the source sends. However, a malicious node can safely drop packets when the probing is not activated, while behaving "normally" when probing is invoked. Hence, the source can never catch the malicious nodes nor bound the malicious dropping rate, unless the probing is always activated which incurs high overhead. In addition, ODSBR employs binary search in the probing phase for dropping localization, until the algorithm converges to a specific link. Since the binary search algorithm proceeds on each packet lost (possibly due to natural loss), in the presence of natural packet loss the algorithm either does not converge or incurs high false positives by incriminating benign links. The security vulnerabilities of Watchers [18, 25], AudIt [10], Fatih [43], and the proposal by Liu et al. [34] were summarized at the beginning of Section 3. The recently proposed Network Confessional [11] also fails to prevent packet modification attacks due to the lack of efficient packet authentication.

High protocol overhead. Among the known secure proposals, the protocol due to Avramopoulos et al. [12] incurs high overhead due to the acknowledgments from all routers in the path and multiple digital signature generation and verification operations for *each* data packet. Both Statistical FL [16] and PAAI-1 [59] achieve small communication overhead, but at the cost of high storage overhead and unacceptably long detection delays (Sections 6 and 7).

Applicability constraints (and security vulnerabilities). A recent proposal due to Wang et al. [53] for forwarding fault localization in sensor networks requires a special tree-like routing infrastructure where the communications take place only between a sensor node and the same trusted base station. Both Watchdog [41] and Catch [40] can identify and isolate malicious routers for wireless ad hoc networks, where a sender S verifies if the next-hop node f_i indeed forwards S's packets by *promiscuously* listening to f_i's transmission. Both approaches rely on the wireless broadcast medium and are thus inapplicable to wired networks. Furthermore, both Watchdog and Catch are vulnerable to collusion attacks, where a malicious node f_m drops the packets of a remote sender S which is out of the promiscuous listening range of f_m while the colluding neighbors in the promiscuous listening range of f_m intentionally do not report the packet dropping behavior of f_m. Finally, TrueNet [60] utilizes trusted computing technologies thus requiring the existence of TPM chips, and is vulnerable to hardware attacks against the TPM chips.

11 Conclusion

In this paper, we design, analyze, implement, and evaluate ShortMAC, an efficient data-plane fault localization protocol, which enables a *theoretically proven* guarantee on data-plane packet delivery and substantially outperforms related protocols in the following aspects. First, ShortMAC achieves high security assurance even in the presence of strong adversaries in control of colluding malicious routers that can drop, modify, inject, and misroute packets at the forwarding paths; whereas a majority of existing fault localization protocols exhibit security vulnerabilities under such a strong adversary model. Second, compared to existing *secure* protocols, ShortMAC achieves several orders of magnitude lower detection delay and protocol overhead, which facilitates its practical deployment. Finally, we demonstrate that ShortMAC's efficient cryptographic operations, even if implemented in software, have negligible effects on the communication throughput via realistic testing on Gigabit Ethernet links. We anticipate that ShortMAC probabilistic authentication and efficient fault localization can become a basic building blocks for the construction of highly secure and efficient network protocols.

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A Proofs for Section 6

A.1 Proof of Lemma 1

Recall from Section 4 that in $\mathrm{ShortMAC}$, the source finds the first C_i^{bad} such that $C_i^{bad} > T_{in}$, and identifies link l_i as malicious. In this proof, we first derive the upper bound β of malicious packet injection (which is based on T_{in}) according to the upper bound δ of false negative rate. Then we calculate the injection threshold T_{in} given the false positive upper bound δ.

With k-bit MACs, when f_{i-1} receives a fake packet, the probability that C_{i-1}^{bad} will be increased is $q = \frac{2^k-1}{2^k}$, since the adversary can only randomly generate a k-bit string for the fake packet without knowledge of the secret keys of other (benign) routers. The probability that C_i^{bad} will be increased is $q(1-q)$.

Malicious Injection Bound. WLOG, suppose f_m is a malicious router and f_{m+1} is benign (there can be other malicious routers between the source and f_m). Suppose the malicious routers between the source and f_m (including f_m) inject y packets on link l_{m+1}. Then whether l_{m+1} will be detected depends on the value of C_{m+1}^{bad}, and the false negative rate $\mathbb{P}_{fn}$ is given by:

$$\begin{aligned}\mathbb{P}_{fn} &= \mathbb{P}(C_{m+1}^{bad} < T_{in}) \\ &= \mathbb{P}\big((q-\epsilon)y < T_{in}\big) \\ &\leq 2e^{-2y\left(q-\frac{T_{in}}{y}\right)^2} \text{(Hoeffding's inequality)},\end{aligned} \tag{5}$$

where ϵ is the deviation and $0 \leq \epsilon \leq q$. To achieve the desired upper bound $\mathbb{P}_{fn} \leq \delta$, we set the threshold β such that

$$2e^{-2\beta(q-\frac{T_{in}}{\beta})^2} = \delta. \tag{6}$$

Solving for β gives:

$$\beta = \frac{T_{in}}{q} + \frac{\sqrt{\left(\ln\frac{2}{\delta}\right)^2 + 8qT_{in}\ln\frac{2}{\delta}} + \ln\frac{2}{\delta}}{4q^2}. \tag{7}$$

(7) implies that if the adversary injects more than β packets on a single link l_{m+1}, C_{m+1}^{bad} will exceed T_{in} and l_{m+1} will be detected with a high probability $\geq 1-\delta$ (or a false negative rate lower than δ).

Injection Detection Threshold. WLOG, suppose f_m is a malicious router and f_{m+1} is benign (there can be other malicious routers between the source and f_m). Suppose the malicious routers between the source and f_m (including f_m) inject y packets on link l_{m+1}. False positives occur when $C_{m+1}^{bad} < T_{in}$ but $C_i^{bad} \geq T_{in}$ (where $i \geq m+2$). (WLOG, suppose f_{i-1} and f_i are honest.) Hence, a benign link l_i is falsely accused, and the false positive rate $\mathbb{P}_{fp}$ is:

$$\begin{aligned}\mathbb{P}_{fp} &:= \sum_{i=m+2}^{d} \mathbb{P}(C_{m+1}^{bad} < T_{in}, C_i^{bad} \geq T_{in} | l_i \text{ benign}) \\ &\leq d \cdot \mathbb{P}(C_{m+1}^{bad} < C_{m+2}^{bad}).\end{aligned} \tag{8}$$

The actual C_{m+1}^{bad} and C_{m+2}^{bad} values can be represented by:

$$\begin{aligned}C_{m+1}^{bad} &= (q-\epsilon_1)\cdot y \\ C_{m+2}^{bad} &= \big(q(1-q)+\epsilon_2\big)\cdot y.\end{aligned} \tag{9}$$

If we can bound

$$\epsilon_1 = \epsilon_2 = \epsilon \leq \frac{p^2}{2}, \tag{10}$$

then we can guarantee that $C_{m+1}^{bad} > C_{m+2}^{bad}$. Therefore, we have:

$$\begin{aligned}\mathbb{P}_{fp} &\leq 1 - \mathbb{P}(\epsilon \leq \frac{q^2}{2}) \\ &= \mathbb{P}(\epsilon > \frac{q^2}{2}) \\ &\leq 2e^{-2y(\frac{q^2}{2})^2}.\end{aligned} \tag{11}$$

Note that in (11), we leverage Hoeffding's inequality and the fact $y \geq T_{in}$ in the false positive cases.

To achieve the desired upper bound $\mathbb{P}_{fp} \leq \delta$, we set the threshold T_{in} such that

$$2e^{-2T_{in}(\frac{q^2}{2})^2} = \delta. \tag{12}$$

Solving for T_{in} gives

$$T_{in} = \frac{2\ln\frac{2d}{\delta}}{q^4}. \tag{13}$$

A.2 Proof of Lemma 2

Drop Detection Threshold and Detection Space. False positives arise when the *observed* drop rate of a benign link l_i, denoted by ρ_i^*, exceeds the drop detection threshold T_{dr}. To bound the total false positive rate below δ, it is sufficient to ensure that each ρ_i^* may exceed T_{dr} with a probability $\delta_i = \frac{\delta}{d}$ (since we need to ensure the overall false positive rate $\sum_i \delta_i \leq \delta$), i.e., $\mathbb{P}(\rho_i^* > T_{dr}) < \frac{\delta}{d}$, which is equivalent to:

$$\mathbb{P}(\rho_i^* - \rho > T_{dr} - \rho) < \frac{\delta}{d}. \tag{14}$$

By using Hoeffding's inequality, we have:

$$\begin{aligned}&\mathbb{P}\Big(\rho_i^* - \rho > T_{dr} - \rho\Big) < 2e^{-2C_{i-1}^{good}(T_{dr}-\rho)^2} \\ &\Rightarrow\ C_{i-1}^{good} \geq \frac{\ln(\frac{2d}{\delta})}{2(T_{dr}-\rho)^2}.\end{aligned} \tag{15}$$

Recall that the check-dropping procedure will detect the malicious link with excessive drop rate closest to the source, denoted by l_m. So we need to guarantee $C_i^{good} \geq \frac{\ln(\frac{2d}{\delta})}{2(T_{dr}-\rho)^2}$ for any $i < m$. Since we also have

$$C_i^{good} \geq N(1-T_{dr})^i \text{ for } i < m, \tag{16}$$

we get:

$$N = \frac{\ln(\frac{2d}{\delta})}{2(T_{dr}-\rho)^2(1-T_{dr})^d}. \tag{17}$$

Analogously, we can also calculate the false negative rate, which yields the same result.

Malicious Dropping Bound. Suppose a malicious node f_m closest to the source receives C_m^{recv} data packets, but claims that it receives C_m^{good} data packets, and drops x fraction of the received C_m^{good} data packets on l_{m+1}. We first have the following facts:

$$\begin{aligned}C_m^{recv} &\leq C_{m-1}^{good} \\ C_{m+1}^{good} &= (1-x)C_m^{recv} + \beta.\end{aligned} \tag{18}$$

To make neither of its incident links undetected, f_m must manage to satisfy:

$$\begin{aligned} \frac{C_m^{good}}{C_{m-1}^{good}} &\geq 1 - T_{dr} \\ \frac{C_{m+1}^{good}}{C_m^{good}} &\geq 1 - T_{dr}, \end{aligned} \tag{19}$$

which yields

$$\begin{aligned} C_{m-1}^{good} &\geq (1 - T_{dr})^{m-1} N \\ &\geq (1 - T_{dr})^d N. \end{aligned} \tag{20}$$

Solving (18), (19) and (20), we have

$$\begin{aligned} x &\leq 1 - (1 - T_{dr})^2 + \frac{\beta}{N(1 - T_{dr})^d} \\ &= \alpha. \end{aligned} \tag{21}$$

A.3 Proof of Theorem 1

$(\alpha, \beta)_\delta-$Statistical Security can directly follow Lemma 1 and Lemma 2. In the following, we will prove $(\Omega, \theta)-$Guaranteed Forwarding Correctness.

Given N and δ, we can set the drop detection threshold T_{dr} from Lemma 2 and the injection bound β from Lemma 1. Let η^{fake} denote the fake data packets the destination has received but not detected yet, and η^{leg} denote the legitimate data packets the destination has received out of N data packets from the source. Then we have:

$$\begin{aligned} \theta &= \frac{\eta^{leg}}{N} \\ &= \frac{C_{d+1}^{good} - \eta^{fake}}{N}. \end{aligned} \tag{22}$$

When no fault is detected in the identify stage, it satisfies:

$$\begin{aligned} &C_{d+1}^{good} \geq (1 - T_{dr})^d N \\ &\eta^{fake} \leq \beta. \end{aligned} \tag{23}$$

By (22) and (23), we have

$$\theta = (1 - T_{dr})^d - \frac{\beta}{N}. \tag{24}$$

Finally, we can integrate $\mathrm{ShortMAC}$ with routing as follows. The control plane first provides a routing path p for the source S, and then avoids faulty links using feedback (fault localization results) from the data plane. In this way, $\mathrm{ShortMAC}$ enables the source to identify the malicious links that reside in previously explored paths. In a network with Ω malicious links, the source can bypass *at least one* of the malicious links after each epoch until a working path is found, resulting in an exploration of at most Ω epochs to find a working path.