The Case for Accelerating BFT Protocols Using In-Network Ordering

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Abstract
Mission critical systems deployed in data centers today are facing more sophisticated failures. Byzantine fault tolerant (BFT) protocols are capable of masking these types of failures, but are rarely deployed due to their performance cost and complexity. In this work, we propose a new approach to designing high performance BFT protocols in data centers. By re-examining the ordering responsibility between the network and the BFT protocol, we advocate a new abstraction offered by the data center network infrastructure. Concretely, we design a new authenticated ordered multicast primitive (AOM) that provides transferable authentication and non-equivocation guarantees. Feasibility of the design is demonstrated by two hardware implementations of AOM— one using HMAC and the other using public key cryptography for authentication — on new-generation programmable switches. We then co-design a new BFT protocol, Matrix, that leverages the guarantees of AOM to eliminate cross-replica coordination and authentication in the common case. Evaluation results show that Matrix outperforms state-of-the-art protocols on both latency and throughput metrics by a wide margin, demonstrating the benefit of our new network ordering abstraction for BFT systems.

1 Introduction
Online services today are commonly deployed in large data centers and rely on fault-tolerance protocols to provide high availability in the presence of failures. An important class of fault-tolerance protocols is state machine replication (SMR). SMR protocols [36, 39, 44, 48, 55] have long been deployed in production systems [18, 20, 24, 32] to ensure a set of distributed nodes behaves like a single, always available state machine, despite failures of individual machines. These protocols, however, can only tolerate node crash failures. In reality, systems running in data centers are facing more sophisticated failures. This is particularly relevant today, as permissioned blockchain systems [3, 8, 50] are increasingly being deployed in data centers for applications such as trading [2, 4]. These systems require tolerance to adversarial nodes and attacks. Recent work [50] has shown that fault tolerance is becoming their main performance bottleneck.

Numerous Byzantine fault tolerant (BFT) protocols [19, 23, 38, 47, 56, 58, 60] have been proposed to handle arbitrary node failures. Their strong failure models, however, come with significant performance implications. BFT protocols typically incur rounds of replica communication coupled with expensive cryptographic operations, resulting in low system throughput and high request latency. To obtain higher throughput, many BFT protocols rely on heavy request batching which leads to long end-to-end decision latency – often in the range of tens of milliseconds. Unfortunately, such latency overheads are prohibitive for modern data center applications with stringent service-level objectives (SLOs). Speculative BFT protocols such as Zyzzyva [38] offer better commitment latency. However, even a single faulty replica would eliminate the performance benefit of these latency-optimized protocols.

In this paper, we propose a new approach to building high-performance BFT protocols in data centers. We observe that traditional BFT protocols are designed with minimum assumptions about the underlying network: The network only provides best effort message delivery. The result of this weak network model is that application-level protocols are responsible for enforcing all correctness properties such as total ordering, durability, and authentication. Our key insight is that strengthening the network model, in which the network can guarantee ordered message delivery, can lead to reduced complexity and performance overhead of BFT protocols. Prior work [42, 43] have shown promises of an in-network sequencing approach for crash fault-tolerant systems. However, existing approaches fail to work once Byzantine failures are introduced: Faulty nodes can disseminate conflicting message orders, and the network sequencer may equivocate by assigning different sequence numbers to each replica. In this work, we propose a new network-level primitive, AOM, that addresses the above challenges. AOM ensures that correct receivers always deliver multicast messages in the same order, even in the presence of Byzantine participants. The key property offered by AOM is transferable in-network authentication.
receivers can verify that a multicast message is properly delivered by AOM, and they can prove authenticity of the message to other receivers in the system. We additionally propose a mixed failure model [45, 54] in which the network infrastructure can be either crash- or Byzantine-faulty. For deployments that trust the network infrastructure, AOM provides ordering guarantees with minimum network-level overhead. For systems that need to tolerate Byzantine network devices, AOM uses a simple round of cross-receiver communication to handle equivocating sequencers.

We demonstrate the feasibility of our approach by implementing our AOM primitive in off-the-shelf programmable switches. The switch data plane performs both sequencing and authentication for AOM messages. While packet sequencing is relatively straightforward, generating secure authentication codes—a complex mathematical procedure—is a major challenge given the switch’s limited resources and computational constraints. We propose two implementations of in-switch message authentication, each with trade-offs among switch resource utilization, performance, and scalability. The first variant implements message authentication code (HMAC) vectors in the switch ASICs using SipHash [9] as the hashing algorithm. The second variant generates signatures using public-key cryptography. Due to hardware constraints, direct in-switch implementation of cryptographic algorithms such as RSA [52] and ECDSA [35] remains infeasible. We design a new heterogeneous switch architecture that tightly couples FPGA-based cryptographic accelerators to the switch pipelines. Our design enables efficient in-network processing and signing of AOM messages, scales linearly with the number of cryptographic accelerators attached, and requires minimum hardware resources on the switch data plane.

Leveraging the strong properties provided by AOM, we co-design a new BFT protocol, Matrix. In the common case, Matrix replicas rely on the ordering guarantees of AOM to commit client requests in a single round trip, eliminating all cross replica communication and authentication. Moreover, Matrix stays in this fast path protocol even in the presence of (up to $f$) faulty replicas, while requiring the theoretical minimum replication factor ($3f+1$). In the rare case of network failures, we design efficient protocols to handle packet drops and faulty switch sequencers while guaranteeing correctness. By evaluating against state-of-the-art BFT protocols, we show that Matrix can improve protocol throughput by up to $3.4 \times$ and end-to-end latency by $42 \times$, demonstrating the benefit of our authenticated in-network ordering approach for BFT systems.

## 2 Background

In this section, we give an overview of state-of-the-arts BFT protocols. We then review recent proposals that use in-network ordering to accelerate crash fault-tolerant systems.

### 2.1 State-of-the-Art BFT Protocols

There has been a long line of work on BFT state machine replication (SMR) protocols. Table 1 presents a summary of the properties and comparison of some recent representative protocols. PBFT [19] is the first practical BFT protocol that tolerates up to $f$ Byzantine nodes, requiring at least $3f+1$ replicas which has been shown to be theoretical lower bound [17]. In PBFT, client requests are committed in five message delays: clients send requests to a primary replica; the primary replica sequences and forwards the requests to the backup replicas; backup replicas authenticate the requests and broadcast their acceptance; after a replica receives quorum acceptance, it broadcasts a commit decision; replicas execute the request and reply to the client once they collect quorum commit decisions. As replicas exchange messages in an all-to-all fashion, each replica processes $O(N)$ messages, making the authenticator complexity $O(N^2)$.

Zyzzyva [38] speculatively executes client requests before they are committed to reduce the communication overhead. The protocol includes a fast path with three message delays when clients receiving matching replies from all replicas, and otherwise a slow path with at least five message delays. The primary replica in Zyzzyva still sends signed messages to all backup replicas ($O(N)$), but with all-to-all communication removed, the authenticator complexity is reduced to $O(N)$. Replicas in Zyzzyva may need to rollback speculatively executed operations during view changes.

Unlike Zyzzyva which pushes the responsibility of collecting authenticators to the client, SBFT [29] uses round-robin message collector among all replicas to remove all-to-all communication, similarly reducing authenticator complexity to $O(N)$. SBFT also leverages threshold signatures to reduce message size, and simultaneously decreases the number of client replies to one per decision.

Many BFT protocols ([19, 29, 38]) use an expensive view change protocol to handle primary leader failure. The standard view change protocol used in PBFT requires $O(N^3)$ message authenticators, limiting the overall scalability. HotStuff [58] addresses this issue by adding an extra phase in normal operation. This design reduces the authenticator complexity of the leader failure protocol to $O(N)$, matching that of the normal case protocol. In return, HotStuff adds a one-way network latency to the request commit delay.

**BFT with trusted components.** To reduce protocol complexity, a recent line of work [23, 25, 41, 56, 61] introduces trusted components on each replica. These trusted components can be implemented in a Trusted Platform Module (TPM) [5] or run in a trusted hypervisor, and are assumed to always behave correctly even if residing on Byzantine nodes.

A2M-PBFT-EA [23] uses an attested append-only memory (A2M) to securely store operations as entries in a log. Each A2M log entry is associated with a monotonically increasing, gap-less sequence number. Once appended, a log entry be-
Table 1: Comparison of Matrix to state-of-the-art BFT protocols. Here, bottleneck complexity denotes the number of messages the bottleneck replica needs to process; authenticator complexity shows the total number of signatures processed by all replicas.

|                        | PBFT [19] | Zyzzyva [38] | SBFT [29] | HotStuff [58] | A2M [23] | MinBFT [56] | Matrix |
|------------------------|-----------|--------------|-----------|---------------|----------|-------------|--------|
| Replication Factor     | 3f + 1    | 3f + 1       | 3f + 1    | 3f + 1        | 2f + 1   | 2f + 1      | 3f + 1 |
| Bottleneck Complexity  | O(N)      | O(N)         | O(N)      | O(N)          | O(N)     | O(N)        | O(1)   |
| Authenticator Complexity| O(N^2)    | O(N)         | O(N)      | O(N)          | O(N^2)   | O(N^2)      | O(N)   |
| Message Delay          | 5         | 3            | 6         | 4             | 5        | 4           | 2      |

3.1 The Case for an Authenticated Ordering Service in Data Center Networks

To guarantee linearizability [30], BFT SMR protocols require all non-faulty replicas to execute client requests in the same order. Due to the best-effort assumption about the network, an application-level protocol is fully responsible for establishing a total order of requests among the replicas. Take PBFT [19] as an example: The primary replica first assigns an order to client requests before broadcasting to backup replicas. All replicas then use two rounds of communication to agree on this ordering while tolerating faulty participants. As discussed in §2, even adding trusted components to each replica does not alleviate the coordination and authentication overhead in BFT protocols: replicas still require remote attestations to verify the received messages.

What if the underlying network can offer stronger guarantees? Prior work [42, 43, 49] have already demonstrated that in-network ordering, realized using network programmability [16, 33], can offer compelling performance benefits to crash fault tolerant SMR protocols. In this work, we argue that BFT protocols can similarly benefit from moving the ordering responsibility to the network. At a high level, by offloading request ordering to the network, BFT replicas avoid explicit communication to establish an execution order, effectively reducing cross-replica coordination and authentication overhead. This network ordering approach improves both protocol throughput – less work is performed on each replica – and latency – fewer message delays are required to commit a request.

However, achieving message ordering in the network is more challenging in a BFT setting. In this section, we discuss the main challenges faced by a BFT-based network ordering primitive and our approach to address them.

Why authenticated ordering in the network? In prior network ordering systems, the network primitive (such as the Ordered Unreliable Multicast in NOPaxos [43] and the multi-sequenced groupcast in Eris [42]) is fully responsibility for assigning an order to client requests. With non-Byzantine participants, this network-level ordering is the only request order observed by any replica. Unfortunately, in a BFT deployment model, a faulty node can easily impersonate the network primitive and assign a conflicting message order, violating the ordering guarantee of the network layer. To tolerate
equivocating Byzantine participants, we augment the network primitive to provide an authentication property: Non-faulty replicas can independently validate that the received message order is indeed established by the network, not by any faulty node. We elaborate on how such authentication can be implemented efficiently on commodity switch hardware in §4.

**Hybrid fault model and Byzantine network.** If the network itself is Byzantine, it can equivocate by assigning different message orders to different replicas, violating the ordering guarantee. In this work, we argue for a dual fault model – one fault model for the participating nodes and one for the network. Our argument is inspired by prior work that proposes hybrid fault model [54] and work [45] that separates machine faults from network faults. Specifically, we always assume a Byzantine failure mode for end-hosts. The network infrastructure, on the other hand, can either be crash-faulty or Byzantine-faulty. The benefit of our approach is that we offer flexibility with an explicit trade-off between fault tolerance and performance. For deployments that trust the network to only exhibit crash and omission faults, our solution offers the optimal performance; if the deployment assumes the network infrastructure can behave arbitrarily, we offer solution that tolerates Byzantine faults in the network, albeit taking a small performance penalty.

We further argue that a hybrid fault model, in which the network is assumed to be crash-faulty, is a practical option for many systems deployed in data centers. Networking hardware presents a smaller attack surface and is less vulnerable to bugs compared to software-based components. They are single application ASICs without sophisticated system software, and formal verification of their hardware designs is common practice. Systems running in data centers also inherently place some trust in the data center hardware infrastructure. It is also against the economic interest of data center operators to violate the trust of their provided services.

Our model resembles existing deployment options in the public cloud: Only deployments that do not trust the cloud infrastructure run their virtual machines on instances with a Trusted Execution Environment (TEE) such as Intel SGX. The majority of use cases place trust on cloud hardware and hypervisors; in return, they attain higher performance compared to their TEE counterpart.

### 3.2 Authenticated Ordered Multicast

We have so far argued for an authenticated ordering service in the network for BFT protocols. To that end, we propose a new Authenticated Ordered Multicast (AOM) primitive as a concrete instance of such model. Similar to other multicast primitives (e.g., IP multicast), an AOM deployment consists of one or multiple AOM groups, each identified by a unique group address. AOM receivers can join and leave an AOM group by contacting a membership service. A sender sends an AOM message to one AOM group, and the network is responsible for routing the message to all group receivers. Note that senders do not know the identity nor the address of individual receivers; they only specify the group address as the destination.

Unlike traditional best-effort IP multicast, AOM provides a set of stronger guarantees, which we formally define here:

- **Asynchrony.** There is no bound on the delivery latency of AOM messages.
- **Unreliability.** There is no guarantee that an AOM message will be received by any receiver in the destination group.
- **Authentication.** A receiver can verify the authenticity of an AOM message, i.e., the message is correctly processed by the AOM network primitive. A correct receiver only delivers authentic AOM messages.
- **Transferable Authentication.** If a receiver $r_1$ forwards an AOM message to another receiver $r_2$, $r_2$ can independently verify the authenticity of the message.
- **Ordering.** For any two authentic AOM messages $m_1$ and $m_2$ that destined to the same AOM group $G$, all correct receivers in $G$ that receive both $m_1$ and $m_2$ deliver them in the same order.
- **Drop Detection.** Receivers can detect AOM message drops in the network and deliver DROP-NOTIFICATION. Formally, for any authentic AOM message $m$, either 1) all correct receivers in the destination group $G$ delivers $m$ or a DROP-NOTIFICATION for $m$ before delivering the next AOM message, or 2) none of the correct receivers in $G$ delivers $m$ or a DROP-NOTIFICATION for $m$.

A key differentiating property of AOM is the capability of receivers to independently verify the authenticity of an AOM message. Authenticity in our case does not refer to the identity of the sender, which still requires end-to-end cryptography. Instead, receivers can verify that a message is correctly processed by the AOM primitive, and the ordering is not forged by other participant in the system. The authentication capability is also transferable: an AOM message can be relayed to any other receiver in the group, and they can independently verify the authenticity of the message.

Another important guarantee of AOM is non-equivocation, i.e., the primitive never delivers conflicting order of messages to non-faulty receivers in a group. As we show in §4, if the network is assumed to be Byzantine-faulty, the primitive requires additional confirmation among the group receivers to guarantee non-equivocation. AOM, however, does not guarantee reliable transmission. The implication is that an AOM message may be delivered by only a subset of the receivers in a group. AOM ensures that receivers who miss the message deliver a DROP-NOTIFICATION. The application-level BFT protocol is then responsible for reaching consensus on the fate of those dropped messages.
4 Design and Implementation of AOM

In this section, we detail our design of the proposed AOM network primitive on reconfigurable switches. Even though ordering and drop detection can be realized using an in-switch sequencing design (§4.2), achieving transferable authentication presents major technical challenges. Here, we present two switch designs that overcome these challenges. The first design implements hash-based message authentication code (HMAC) vectors directly in the switch ASICs (§4.3). The second design proposes a new switch architecture that integrates cryptographic accelerators with the switch pipelines. This architecture enables us to sign AOM messages using public key cryptography in the network (§4.4).

4.1 Design Overview

Our AOM primitive design consists of three major components: a network-wide configuration service, a programmable network data plane, and an application-level library running on AOM senders and receivers. Similar to IP multicast, receivers create and join an AOM group by contacting the configuration service through secure TLS channels. The configuration service selects one programmable switch in the network as a sequencer for the group, and instructs the sequencer switch to broadcast routing advertisement for the group address (e.g., using BGP). Once the advertisement is propagated, AOM messages destined to the group address will be forwarded to the sequencer switch. The switch implements sequencing (§4.2) and authentication (§4.3, §4.4) of AOM packets in the data plane using the P4 language [16]. When sending an AOM packet, the sender-side library generates a custom packet header (after the UDP header) that contains the group ID, a sequence number, an epoch number, and an authenticator. Except the group ID, all remaining fields in the custom header are filled by the sequencer switch. The receiver-side library verifies the authenticator and delivers AOM messages in sequence number order. For any gap in the number sequence, the receiver delivers a DROP-NOTIFICATION. For deployments that assume a Byzantine-faulty network, the receiver-side library additionally exchanges ORDER-CONFIRM messages with other receivers in the group to tolerate sequencer equivocation (§4.2).

4.2 Message Ordering, Drop Detection, and Failure Handling

To establish a consistent ordering of AOM messages, we leverage programmable switches to stamp monotonically increasing, gap-less sequence numbers to each AOM packet. The sequencer switch maintains a counter in a data plane register for each AOM group. When processing a AOM packet, the switch uses the destination group address to locate the group counter register, increments the counter, and writes the counter value into the message header. After the switch generates an authenticator for the packet, it uses its replication engine to multicast the stamped AOM message to all AOM receivers in the group. As discussed in §4.1, receivers deliver authenticated AOM messages in sequence number order.

Our sequencing approach enables drop detection: When a receiver notices a gap in the message number sequence — indicating some AOM messages have been dropped in the network — it delivers DROP-NOTIFICATIONS to the application. This naive approach still works when messages are reordered in the network: A receiver simply ignores out-of-order messages. However, it can lead to more frequent DROP-NOTIFICATIONS when packet reordering rate is high. As an optimization, receivers buffer out-of-order AOM messages and only deliver DROP-NOTIFICATION after a timeout.

Tolerating Byzantine-faulty network. If the network infrastructure is non-Byzantine (§3.1), a receiver can directly deliver authenticated AOM messages in the sequence number order. This delivery rule satisfies our ordering property: Any two receivers are guaranteed to receive identical messages for each sequence number. If a Byzantine-faulty network is assumed, the sequencer may equivocate by sending different request ordering to each receiver. To tolerate sequencer equivocation, for each received AOM message, a receiver broadcasts a signed confirmation that includes the sequence number and the message hash. It ignores subsequent AOM messages with the same sequence number. A receiver only delivers an AOM message after it collects enough matching confirmations (at least \(2f + 1\) where \(f\) is the number of faulty receivers) for the message. This strengthened delivery rules ensures our ordering property in a Byzantine-faulty network, since no two non-faulty replicas can deliver distinct AOM messages for the same sequence number. When delivering, the receiver-side library delivers both the AOM message and the collection of confirmation messages. This entire message set serves as an ordering certificate — it can be used to prove the correct ordering of the associated AOM message. On the other hand, a single AOM message alone can serve as the ordering certificate in a crash-faulty network.

Sequencer switch failover. The above scheme only ensures message ordering and drop detection when there is a single switch sequencing AOM messages. To tolerate faulty switches, AOM receivers request the configuration service to fail over to a different sequencer for the group. However, receivers in a group may not deliver the same set of messages from the failed sequencer when the new sequencer switch starts. Furthermore, the previous sequencer may only suffer a transient fault; it may continue to process AOM messages concurrently with the new sequencer. To properly handle a switch failover, the application-level protocol is responsible for reaching consensus on the set of messages delivered by the failed sequencer (§5.5). Once an agreement is reached, the receivers ask the configuration service to choose a new sequencer switch, and exchange the necessary authentication keys. They then start delivering AOM messages and DROP-NOTIFICATIONS for the new sequencer switch, ignoring mes-
sages from the old switch.

4.3 HMAC-Based In-Network Authentication

Compared to message sequencing, generating secure and transferable authentication tokens for messages directly in network hardware raises bigger technical challenges. Common authentication mechanisms such as Hash-based Message Authentication (HMAC) and public key cryptography involve long sequence of complex mathematical operations, which are challenging tasks for the resource-constrained network devices. In this work, we demonstrate the feasibility of implementing in-network authentication on current generation programmable switches.

Our first design uses HMAC vector as the authentication token. Specifically, when a receiver joins a AOM group, it uses a key exchange protocol [46] to share a secret key with the current sequencer switch, facilitated by the configuration service. The switch control plane installs the secret key of each receiver in the data plane. When sending AOM messages, the sender generates a digest of the message using a collision-resistant hash functions [51], and adds the digest to the message header. The switch concatenates the stamp sequence number §4.2 and the message digest, and generates a vector of HMAC hashes, one for each receiver. For each hash, the switch uses the digest of the message, the sequencer number, the secret key of the receiver, and a cryptographic hash function as the input. It then writes the entire HMAC vector into the message header. A receiver authenticates a AOM message by checking if a locally computed HMAC hash matches the corresponding entry in the received HMAC vector.

In-network HMAC implementation. Implementing an unforgeable HMAC requires access to collision resistant cryptographic hash functions. Common options such as MD5 [51] and SHA-2 [27] either require modular arithmetic or rounds of sequential heavy computations, making them impractical to implement on programmable switches. Recent advancements – HalfSipHash [59] and P4-AES [22] – demonstrate the feasibility of switch implementation of cryptographic hash functions with throughput up to ~150 million hashes per second [59]. We therefore leverage HalfSipHash as a building block for our in-switch HMAC design.

A cryptographic hash function, unfortunately, only solves half of the equation. Implementing a HMAC vector in the switch data plane still requires us to overcome several technical challenges. Firstly, HalfSipHash consumes non-negligible switch hardware resources (e.g., using all 12 pipeline stages), making parallel generation of HMAC vector difficult. With careful optimizations, we managed to fit four parallel instances of HalfSipHash within a single switching pipeline. We explicitly trade-off the number of pipeline passes for HMAC computation to achieve higher degree of parallelism, while maintaining high throughput (§6.1).

Second, there exists data dependencies between HMAC computation and other packet processing logic such as sequencing. A naive combination of the logic would result in a dependency chain exceeding the hardware limit [37]. To address the issue, we perform pipeline-folding (Figure 1a) to extend our computation beyond the available pipeline stages. Furthermore, our approach decouples HMAC vector computation from other packet processing logic, resulting in a simpler and more modular design.

Lastly, the number of HMAC instances grows linearly with the multicast group size. By dedicating a pipeline for HMAC vector computation, our design can scale up by configuring additional ports in the HMAC pipeline as loopback ports. We arrange receivers into logical groups of four, and perform HMAC computations for the different groups in parallel. Receiver scalability is then only bounded by the configurable loopback ports (up to 72 receivers in our current design).

Putting things together, we illustrate our design in Figure 1a. We dedicate switch PIPE 1 for HMAC vector computation, while the senders and receivers are connected to PIPE 0. When an AOM packet enters ingress of PIPE 0, the switch sequences the packet and performs the necessary pre-processing (e.g., header initialization) for HMAC computation. The packet is then multicasted to the dedicated loopback ports to perform HMAC computations in parallel. Each HMAC computation requires 12 pipeline passes in PIPE 1. The resulting packet is then multicasted and forwarded to all receivers after post-processing (e.g., copying the HMAC result) at PIPE 0 egress.

4.4 In-Network Public-Key Cryptography

One issue with our HMAC vector design is scalability: the number of HMAC instances and the vector size increases linearly with the multicast group size. The amount of switch resources and the packet header size thus bound the deployment scale. For deployments that demand bigger group sizes, we propose a second design that implements message authentication using public-key cryptography. Concretely, each sequencer switch maintains a secret private key, and a public key which it shares with the configuration service. The service forwards the switch’s public key to a receiver when it joins the AOM group. When processing AOM messages, the switch produces a digital signature on the concatenated message digest and sequence number using a public key algorithm [35] and its private key. It then writes the signature to the message header, before multicasting to all group receivers. Receivers use the switch public key to verify the authenticity of the message. Our in-network public-key approach is agnostic to the number of receivers: the switch only maintains one private key, and it writes a single signature to an AOM message, regardless of the number of receivers.

In-network cryptography design. Implementing public-key cryptography in a network switch is a daunting task. The

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1 We extend P4-AES to implement AES-CMAC [34]. However, for brevity yet without the loss of generality, we only focus on HalfSipHash for the remainder of the discussion.
RSA [52] public-key algorithm requires modular exponentiation of large prime numbers. Even with aggressive optimizations, calculating a signature still involves unbounded loops of multiplications and modulo operations. The ECDSA [35] algorithm involves similar complexity, and additionally requires random number generations and multiplicative inverses. None of these operations are supported on current generation programmable switches. Due to strict timing, power, and resource constraints, future programmable switches are unlikely to add support for these computations.

To circumvent the limitations of existing switches, we propose a new switch architecture that adds a specialized cryptographic co-processor alongside the main switching chip, as shown in Figure 1b. The co-processor includes a simple processing element, dedicated fast memory, and one or more cryptographic accelerators. It is connected to the switching chip through high-speed PCIe, interconnect (such as QuickPath Interconnect), or internal network links. This architecture is highly feasible: PCIe links between the switching chip and the control plane CPU is commonplace in off-the-shelf switches, and switch vendors have roll-outed commercial products that couple powerful FPGAs with the switch ASIC [1]. When the switch determines that a packet requires a cryptography operation, it writes the operation input (e.g., operation type and crypto key identifier) into pipeline metadata. After egress pipeline processing, the switch submits both the metadata and the packet to the co-processor who performs the crypto operation. The co-processor then writes the result into the header, and sends both the packet and the metadata back to the switch for further processing.

**Cryptographic accelerator implementation.** We implement an accelerator prototype that performs AOM signature signing on a Xilinx Alveo U50 FPGA card [7]. Figure 1c shows the high-level architecture of our hardware design. The card is connected to one of the switch ports through a 100Gbps QSFP28 cable. After an incoming packet is parsed by a parser module, a SHA-256 [27] hashing module calculates a hash of the concatenated message digest and the stamped sequence number. A signing module then generates a signature of the hash using the secp256k1 elliptic curve [35]. To reduce signing latency, part of the computation is done by a pre-compute module before the hash is received by the signer. Finally, a stream merger module sends the signed packet back to the QSFP28 port. All of the hardware modules, except the Xilinx QSFP28 hard IP, are developed in-house using a mixture of RTL and HLS.

To handle message rate higher than the signature generation throughput, we apply a hash chaining technique. Each AOM packet is additionally stamped with a hash of the preceding packet in the number sequence. For packets that do not have a signature, receivers wait until the next signed packet, and verify the entire batch by checking the chain in the reverse order. This chaining optimization is implemented in the hardware hashing module. Additionally, we design a signing ratio controller to govern the frequency of signature generation based on traffic rate and signing throughput.

**5 The Matrix Protocol**

Leveraging the authenticated ordering guarantee provided by our AOM network primitive, we co-designed a new BFT protocol, Matrix, that commits client operations in a single RTT, even in the presence of Byzantine replicas. In the rare case of packet drops in the network, Matrix takes advantages of AOM’s drop detection to efficiently recover from missing operations. To tolerate faulty replicas and network sequencers, Matrix runs a unified view change protocol to replace the faulty participants.

**5.1 System Model**

We assumes a Byzantine failure model for clients and replicas of the protocol. Each Matrix instance is assigned a unique AOM group address. When a replica joins a Matrix instance, it registers as a group receiver of the corresponding AOM group. Besides regular asynchronous unicast messages, nodes in the system can send AOM messages to a Matrix group.

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**Figure 1:** Figure 1a shows the folded pipeline design for implementing HMAC on a switch. Figure 1b illustrates our new switch architecture that tightly couples a crypto accelerator with the switch pipeline. Blue arrows denote unsigned AOM packets whereas the red arrows represents signed AOM packets. The thick arrows refers to multicast. Figure 1c shows the FPGA hardware design of our crypto accelerator.
AOM messages follow the properties as defined in §3.2. We make standard cryptography assumptions: Nodes do not have enough computational resources to subvert the cryptographic hash functions, message authentication codes, and public-key cryptography algorithms we used in the protocol. We also assume a strong adversary model: Byzantine nodes can collude with each other, but they cannot delay correct nodes indefinitely.

Matrix is a state machine replication protocol [53] protocol. We assume all operations executed by the protocol are deterministic. With less than $\lfloor \frac{n}{3} \rfloor$ (Byzantine) faulty replicas in the system (where $n$ is the total number of replicas), Matrix guarantees linearizability [30] of client operations. Due to the impossibility of asynchronous consensus [28], Matrix only ensures liveness when the system is in period of synchrony.

5.2 Protocol Overview

Matrix relies on the guarantees provided by the AOM network primitive to achieve single RTT commitment in the common case. Specifically, clients multicast requests to Matrix replicas. In the absence of network-level anomalies (e.g., message drops and switch failures), all replicas delivers AOM messages in the exact same order. Crucially, such guarantee implies that replicas require no explicit communication to agree on the order of messages. Matrix thus avoids the expensive cross-replica coordination and server signature signing/verification required by other BFT protocols. Moreover, adversaries can not temper with the order of messages nor their content, as correct replicas can independently verify the authenticity and integrity of each AOM message. Once an AOM message is delivered, replicas can immediately execute the request and respond to the client, resulting in a single phase fast path protocol. As discussed in §4.2, when delivering messages, the network primitive provides an ordering certificate: the single AOM message for a crash-faulty network, or the additional 2$f+1$ confirmations for a Byzantine-faulty network. Similar to previous speculative protocols [38, 43, 49], Matrix relies on clients to confirm operation durability. However, since all correct replicas already established a total order of operations, Matrix does not require extra protocols to handle faulty replicas (Zyzzyva [38]) or state divergence due to out-of-order speculative executions (Speculative Paxos [49]).

In the rare case where AOM messages are dropped in the network, the AOM primitive delivers DROP-NOTIFICATIONS to non-faulty replicas. To handle DROP-NOTIFICATIONS, replicas only need to agree on whether to process or to skip the message, not the order of messages. Matrix uses a BFT binary consensus protocol, driven by a leader replica, to reach this agreement. In this protocol, the leader uses a single ordering certificate (received by any replica) to commit the corresponding message. To permanently skip the message, the leader replica collects evidences from a quorum of replicas to form a drop certificate, which non-leader replicas would verify before committing the message as a NO-OP. We use drop certificates to prevent Byzantine replicas from delaying the agreement indefinitely.

A faulty AOM sequencer may stop multicasting messages or deliberately equivocating or dropping messages. To ensure progress, Matrix replicas request the network to replace the faulty sequencer. Installment of a new sequencer indicates the start of a new epoch. Correctness of the protocol requires replicas to agree on the set of messages processed in the last epoch before entering the new epoch. To that end, each Matrix instance goes through a sequence of views: each view is identified by a view number represented as a $(\text{epoch-num}, \text{leader-num})$ 2-tuple. We establish a total order – defined by the lexicographical order of the 2-tuple – on the set of view numbers. When the current leader replica has failed (or suspected to be failed) or an old epoch has ended, replicas advance the respective field in the view number, and use a view change protocol [19, 43, 44, 49] to reach agreement on the set of messages in the last view.

Matrix also includes a synchronization protocol that periodically synchronizes replica states and commits speculatively executed requests. Details of the synchronization protocol, as well as formal safety and liveness proofs of Matrix, can be found in §A.

5.3 Normal Operation

We first consider the common case protocol in which AOM messages, instead of DROP-NOTIFICATIONS, are delivered to Matrix replicas in a stable epoch. A client $c$ requests execution of an operation $op$ by sending a signed message $\langle \text{REQUEST}, op, request-id \rangle_\sigma$, using the AOM primitive, where request-id is a client-generated identification to match replica replies. The message is processed by the network primitive (§3.2), and an ordering certificate (OC) for the message is delivered to all replicas. If the client does not receive replies in a timely manner, it uses regular unicast to send the request to all replicas directly (while keeps resending the request using AOM). Our view change protocol (§5.5) ensures that the request is eventually committed even if the AOM sequencer is faulty.

Replica $i$ verifies the OC by authenticating the AOM authenticator and checking the $2f+1$ matching confirmations (only for Byzantine-faulty network). It then adds the OC to its log, speculatively executes $op$, signs and replies $\langle \text{REPLY}, \text{view-id}, i, \text{log-slot-num}, \text{log-hash}, \text{request-id}, \text{result} \rangle_\sigma$ to the client, where view-id is the current view number, log-slot-num is the log index the request occupies, log-hash is a hash of the log up to the index, and result is the execution result. Similar to prior work [49], we use hash chaining for $O(1)$ hash calculation.

Client $c$ waits for $2f+1$ replies from different replicas with valid signatures and matching view-id, log-slot-num, log-hash, and result. It then accepts the result in the reply.

5.4 Handling Dropped Messages

When a non-leader replica $i$ receives a DROP-NOTIFICATION, it attempts to recover the missing message from the leader.
To do so, it sends a \( \langle \text{QUERY}, \text{view-id}, \text{log-slot-num} \rangle \) to the leader. \text{QUERY} messages require no signatures since they do not alter the state of a correct replica. If the leader has the corresponding \text{OC}, it responds with a \( \langle \text{QUERY-REPLY}, \text{view-id}, \text{log-slot-num}, \text{OC} \rangle \). Replica \( i \) verifies the \text{OC} and ensures the enclosed \text{AOM} message is the missing message by checking the internal sequence number. It then resumes normal operation. Because \text{OC} can be independently verified by any replica, \text{QUERY-REPLYS} also require no signatures. Note that replica \( i \) blocks on waiting for the leader’s response or a committed \text{NO-OP} before processing subsequent client requests, resending \text{QUERY} messages if necessary.

If the leader \( l \) itself receives a \text{DROP-NOTIFICATION}, it broadcasts a \( \langle \text{GAP-FIND-MESSAGE}, \text{view-id}, \text{log-slot-num} \rangle \) to all replicas. When replica \( i \) receives a \text{GAP-FIND-MESSAGE}, it replies to the leader with either a \( \langle \text{GAP-RECV-MESSAGE}, \text{view-id}, \text{log-slot-num}, \text{OC} \rangle \) if it has received the ordering certificate, or a \( \langle \text{GAP-DROP-MESSAGE}, \text{view-id}, i, \text{log-slot-num} \rangle \) if it has also received a \text{DROP-NOTIFICATION} for the message. If a replica replies \text{GAP-DROP-MESSAGE} to a \text{GAP-FIND-MESSAGE}, it blocks until it receives the gap agreement decision (ignoring \text{QUERY-REPLYS} for the message).

Once the leader receives one \text{GAP-RECV-MESSAGE} or \( 2f + 1 \) \text{GAP-DROP-MESSAGE} (including from itself), whichever happens first, it uses a \text{binary Byzantine agreement protocol}, similar to PBFT [19], to commit the decision. Replicas store the quorum prepare (\text{GAP-PREPARE}) and commit (\text{GAP-COMMIT}) messages it collects during the agreement protocol for future view changes. We refer to the quorum \text{GAP-COMMITS} as a \text{gap certificate}. Once the decision is committed, replicas store either the \text{OC} (message is received) or a \text{NO-OP} (message is permanently dropped) in their logs, and resume normal operation. Details of this agreement protocol can be found in §A.1.

5.5 View Changes

We use a view change protocol, inspired by PBFT, to handle both leader failures and faulty \text{AOM} sequencers. The protocol guarantees that all committed operations (including \text{NO-OPs}) will carry over to the new view. View changes can be initiated in two scenarios: (i) when a non-leader replica in view \( (e, l) \) fails to make progress in a gap agreement or state synchronization protocol after a timeout, it initiates a view change with a new view \( (e, l + 1) \), and (ii) when a replica receives a request message directly from the client (§5.3) but the request is not delivered by \text{AOM} after a timeout, it initiates a view change with a new view \( (e + 1, l) \).

For view changes that involves switching epochs, the protocol requires log consistency before entering the new epoch. To that end, we introduce an \text{epoch certificate} consisting of \( 2f + 1 \) valid \text{EPOCH-START} messages from distinct replicas. An epoch certificate is a proof of the agreed starting log position of the epoch. We then define \text{validity} of a replica log as the following: a replica log is valid if and only if (i) the starting log position of all epochs are supported by a valid \text{epoch-cert}, and (ii) within each epoch \( e \), all log positions are filled with either a valid \text{OC} or a \text{NO-OP} supported by a gap certificate.

Our view change protocol is similar to the one in PBFT. The main differences are the additional epoch certificates and the definition of log validity. Details of the protocol can be found in §A.2.

5.6 Correctness

Here, we sketch a proof of correctness for the Matrix protocol. Complete safety and liveness proofs can be found in §B.

The safety property we are proving is linearizability [30]. A key definition we use in our proof is \text{committed operations}: an operation is committed in a log slot if it is executed by \( 2f + 1 \) replicas with matching \text{view-ids} and \text{log-hashes}.

First, we show that within an epoch, if a request \( r \) is committed at log slot \( l \), no other request \( r' (r' \neq r) \) or \text{NO-OP} can be committed at \( l \). Due to the guarantees of \text{AOM}, no correct replicas will execute \( r' \) at log slot \( l \) given that some correct replica has already executed \( r \), so \( r' \) can never be committed at \( l \). To show that \text{NO-OP} cannot be committed, we prove by contradiction. Assume a \text{NO-OP} is committed at \( l \), some replica would have received a \text{GAP-DECISION} with \( 2f + 1 \) valid \text{GAP-DROP-MESSAGE} from the leader, and \( 2f \) \text{GAP-PREPARE} containing a \text{drop} decision from different replicas. By quorum intersection, no replica can receive \( 2f \) distinct \text{GAP-PREPAREs} with a \text{recv} at \( l \), making \text{drop} the only possible decision. Our view change protocol also ensures that the decision will persist in all subsequent views. Moreover, \( 2f + 1 \) replicas have sent a \text{GAP-DROP-MESSAGE}, and at least \( f + 1 \) of those replicas are non-faulty. Since they block until they receive \text{GAP-COMMITS} and the only possible outcome is \text{drop}, they will not execute \( r \). By quorum intersection, \( r \) cannot be committed, leading to a contradiction.

Next, we show that within an epoch, if a request is committed at log slot \( l \), all log slots before \( l \) will also be committed. A committed request at \( l \) implies that at least \( f + 1 \) correct replicas have matching logs up to \( l \). For each log slot before \( l \), since the same request has been processed by \( f + 1 \) correct replicas, we can apply the same reasoning as above to show that no other request or a \text{NO-OP} can ever be committed at that slot.

Lastly, we show that our view change protocol guarantees that correct replicas agree on all committed requests and \text{NO-OPs} across views, and that they start each epoch in a consistent log state. The first point is easy to show given that our view change protocol merges \( 2f + 1 \) logs and using the quorum intersection principle. To prove the second point, we only need to show that for each epoch \( e \), all correct replicas end \( e \) at the same log slot before starting \( e + 1 \). To enter epoch \( e + 1 \), a correct replica needs a valid \text{epoch-cert} for epoch \( e + 1: 2f + 1 \) distinct \text{EPOCH-STARTs} with matching \text{log-slot-num}. 
By quorum intersection, no other epoch-cert can exist for $e + 1$ with a different log-slot-num. And since correct replicas verify epoch-cert for every epoch during view changes, by induction, their logs will be in consistent state.

6 Evaluation

We implement Matrix and the AOM library in $\sim$1600 lines of Rust code. The HMAC version of the AOM sequencer is implemented in $\sim$1900 lines of P4 [16] code compiled using the Intel P4 Studio version 9.7.0. The FPGA-based cryptographic accelerator comprises of $\sim$1500 lines of HLS C++/Verilog code, synthesized using the Xilinx Vivado Design Suite 2020.2 [6].

We compare Matrix against PBFT [19], HotStuff [58], Zyzzyva [38]. For a fair comparison, all protocols are implemented in the same Rust-based framework. We implemented the event-driven version of HotStuff, which closely matches their open-source implementation [31]. We also added batching support to PBFT, HotStuff, and Zyzzyva, following the respective batching technique proposed in the original work. Matrix does not use any batching on the protocol level. Only when using the public key cryptography variant of AOM, Matrix replicas buffer packets until receiving a signed message from the network. Unless specified otherwise, we ran all protocols on four replicas, tolerating one Byzantine failure. As a baseline, we also implemented an unreplicated protocol that runs on one server, tolerating no failures.

Testbed. Our testbed consists of nine servers and a Xilinx Alveo U50 FPGA, all connected to an Intel Tofino-based [33] programmable switch. Replicas ran on machines with dual 2.90GHz Intel Xeon Gold 6226R processors (32 physical cores), 256 GB RAM, and Mellanox CX-5 EN 100 Gbps NICs. Clients ran on a machine with one 2.10GHz Intel Xeon Gold 6230 processor (20 physical cores), 96 GB RAM, and Mellanox CX-5 EN 100 Gbps NICs.

6.1 Micro-benchmarks

We first conduct micro-benchmarks to measure the performance and resource usage of our AOM network primitive. We evaluate both the HMAC-based switch pipeline and the FPGA-based cryptographic accelerator.

HMAC-based in-switch authentication. Table 2 shows the switch resource usage of our in-network HMAC vector design §4.3. To measure the latency and sustainable throughput of our design, we generate 64 byte AOM packets at line rate using the Tofino built-in packet generator. The average latency for generating one HMAC vector (size of four) is 8.5μs. Our switch design attains a maximum HMAC throughput of 5.7Mpps, which is around 22 million hashes per second.

FPGA-based ECDSA accelerator. Hardware resource usage of our crypto accelerator (§4.4) is summarized in Table 3. Using RTL-level simulation at 230MHz clock rate, latency of packet processing without signature signing is 835ns, giving a maximum throughput of 1.20M packets/sec. The secp256k1 signer module incurs around 745ns additional latency. The burst signing rate is at 57M signatures/sec, while the sustained signing rate is around 81.78K signatures/sec, limited by the pre-calculation module.

We also measured end-to-end performance by generating AOM packets from an end-host to the FPGA through the Tofino switch. To accurately measure accelerator latency, the switch adds hardware timestamps before forwarding to the accelerator and after receiving the reply. Signing latency is measured at 2.36μs (including the Ethernet module). The maximum overall throughput is 1.20M packets/sec, and the sustainable signing throughput is 71.7K signatures/sec.

6.2 Latency vs. Throughput

We next evaluate the latency and throughput of Matrix and compare them to other BFT protocols. Our focus here is protocol-level performance, so we ran a simple echo-RPC...
application: client requests are randomly generated strings, and the state machine running on the replicas simply echo the same string back to the client. We used an increasing number of closed-loop clients, and measured the end-to-end latency and throughput observed by the clients.

As shown in Figure 2, HMAC-based Matrix achieves higher maximum throughput than PBFT (2.5×) and HotStuff (3.4×). More aggressive batching can further increase HotStuff’s throughput to a level comparable to Matrix; however, its latency also increases to more than 10ms. To commit a client operation, these protocols require explicit coordination among the replicas, with each message requiring expensive cryptographic operations. Matrix, on the other hand, leverages guarantees of AOM to eliminate coordination and cross-replica authentication overhead in the common case. Comparing to Zyzzyva, Matrix still achieves 1.8× higher throughput. Moreover, when one of the replicas becomes faulty, throughput of Zyzzyva drops by more than 54%, while throughput of Matrix is unaffected. When using the public-key variant of AOM, Matrix only suffers a 60K throughput decrease, despite requiring more expensive cryptographic operations. It demonstrates the efficiency of our in-network crypto accelerator design.

Figure 2 also shows the bigger benefit of Matrix— latency. HMAC-based Matrix outperforms PBFT in latency by 14.68×, HotStuff by 42.28×, and Zyzzyva by 8.56×. Matrix commits client operations in two message delays, while the other three protocols require at least three message delays with additional authentication penalties. Using the public-key variant of AOM adds about 55µs to the latency of Matrix. However, this version of Matrix still outperforms all the other protocols in latency by at least 2.7×.

Tolerating Byzantine network. As discussed in §4.2, to tolerate Byzantine network sequencers, AOM receivers exchange and authenticate confirmation messages. This can lead to degraded throughput and latency compared to the non-Byzantine network variant. Figure 2 shows the performance of Matrix when tolerating a Byzantine network. By batch processing of confirmation messages, Matrix minimizes the impact of the additional message exchanges, and is able to sustain a high throughput at the expense of higher latency. As shown in the figure, this Matrix variant still outperforms the other comparison protocols in both throughput and latency.

6.3 Protocol Scalability
To evaluate the scalability of Matrix, we gradually increase the number of Matrix replicas (i.e., capable of tolerating more Byzantine failures), and measure the maximum sustainable throughput. Due to the limited number of physical servers in our cluster, we deploy up to 13 replicas, with ƒ replicas not running to simulate faulty nodes. As shown in Figure 3, Matrix is able to scale to 13 replicas with only a ∼2.39% throughput drop. Matrix replicas process a constant number of messages per client request, regardless of the replica count. This allows Matrix to scale its performance almost linearly with more replicas. Adding replicas, however, would increase the number of reply messages Matrix clients need to receive. Matrix effectively shift the collector load to the client, which can naturally scale.

6.4 Resilience to Network Anomalies
When AOM messages are dropped in the network, Matrix replicas coordinate to agree on the fate of the message. To evaluate Matrix’s resilience to network anomalies, we simulate packet drops in the network, and measure the maximum throughput of Matrix. As shown in Figure 4, throughput of Matrix is largely unaffected when moderate amount of packets are dropped. This is due to the drop detection property of AOM: non-faulty Matrix replicas can efficiently recover missing messages from each other, without the expensive agreement protocol. When a higher percentage of packets are dropped (1%), Matrix does suffer a more observable throughput drop.

Sequencer switch failover. When the sequencer switch becomes faulty, Matrix performs a view change and fails over to a different sequencer. To understand the impact of a faulty sequencer, we measure the throughput of Matrix during a switch failover. Throughput of Matrix dropped to zero after the sequencer became faulty. After replicas detect the fault, they ran a view change protocol which finished in less than 200
Max Throughput (txns/sec)

0K 50K 100K 150K 200K 250K

Unreplicated Neo-HM Neo-PK Neo-BN Zyzzyva ZyzzyvaFPBFT HotStuff

Figure 5: Comparing performance of replicated key-value store using YCSB

µs. They then inform the network controller to pick a new sequencer. After rerouting is done (around 100 ms), throughput of Matrix quickly resumed to its peak.

6.5 BFT Storage System Performance

Lastly, we evaluated the performance of Matrix when running more complex real-world applications, and compared against other protocols. We developed an in-memory, B-Tree-based key-value store, and ran YCSB workload A with 100K records and 128-bytes fields. Maximum YCSB throughput attained by each system is shown in Figure 5. Matrix achieved higher throughput than PBFT, HotStuff, and Zyzzyva when running a more complex application. This application requires protocols to handle larger requests than previous experiments, which reduces the efficiency of batching for Zyzzyva, PBFT and HotStuff. Matrix exploits its lower message complexity to attain higher performance.

7 Related Work

BFT protocols. As discussed in §2.1, there has been a long line of work on designing practical BFT protocols [19, 29, 38, 58]. These protocols guarantee correctness in an asynchronous network, and ensure liveness during weak asynchrony. They all use a single leader node to coordinate ordering and agreement, and rely on view change (or similar) protocols to deal with faulty leaders. Byzantine Paxos [40] and DBFT [26] propose a leaderless BFT design, but require a synchronous protocol that commits in $O(f)$ or more rounds. HoneyBadger [47] attacks the weak synchrony assumption and provides optimal asymptotic efficiency. However, it introduces $O(N^3)$ message complexity and five message delays. Matrix leverages the guarantees of AOM to eliminate the leader and coordination overhead in the common case, leading to a bottleneck message complexity of $O(1)$ and two message delays to commit an operation.

BFT with trusted components. Recent work have proposed leveraging trusted components to improve BFT protocols [12, 23, 25, 41, 56]. Using a local trusted component on each replica enables these protocols to reduce the replication factor from $3f + 1$ to $2f + 1$. However, since the trusted components are all local to replicas, they still necessitate coordination among replicas to commit each client operation. Moreover, many of these protocols propose to implement trusted components in resource-constrained TPM hardware which significantly limit their performance. Matrix implements its ordering service in the data center network. Relying on authenticated network ordering, the protocol avoid all coordination in normal operation. And by implementing on fast networking hardware, the service does not become the performance bottleneck.

Network ordering. A classic line of work in distributed computing proposes stronger network models, such as atomic broadcast [15, 36] and virtual synchrony [13, 14], to simplify distributed system designs. These network primitives guarantee that a total order of messages are delivered to all broadcast receivers. However, atomic broadcast and virtual synchrony do not offer performance benefits to distributed systems – implementing them is equivalent to solving consensus [21]. NOPaxos [43] and Eris [42] pioneered a weaker network model in which messages are delivered in a consistent order but reliable transmission is not guaranteed. This weaker model can be efficiently implemented using programmable switches and other networking hardware. NOPaxos proposes a Ordered Unreliable Multicast primitive for state machine replication, while Eris designs a multi-sequenced groupcast primitive for distributed transactions. BIDL [50] uses sequencers to parallelize consensus and transaction execution in a permissioned blockchain system. However, BIDL still uses traditional BFT protocols for consensus and its sequencer design does not improve performance of the BFT protocol itself. The network sequencing approach has also been applied to other distributed system designs [10, 11, 57]. Our AOM primitive was inspired from these work. Moreover, AOM provides transferable authentication guarantee, which is crucial for BFT protocols.

8 Conclusion

In this work, we presented a new approach to designing high performance BFT protocols in data centers. Our work proposed a novel in-network authenticated ordering service. We demonstrated the feasibility of this design by implementing two network primitive variants, one using HMAC and another using public key cryptography for authentication. We then co-designed a new BFT SMR protocol, Matrix, that eliminate cross-replica coordination and authentication in the common case. Matrix outperforms state-of-the-art BFT protocols on both throughput and latency metrics.

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A Additional Protocol Details

A.1 Gap Agreement Protocol

When the leader in Matrix receives a DROP-NOTIFICATION, it broadcasts a GAP-FIND-MESSAGE, and waits for either a single GAP-RECV-MESSAGE or 2f + 1 GAP-DROP-MESSAGES. It then uses a binary Byzantine agreement protocol to commit the decision. In this section, we describe the detail of the agreement protocol, which is omitted in the main paper.

The leader broadcasts a \langle GAP-DECISION, view-id, log-slot-num, decision\rangle_{\sigma}, where decision is either a single GAP-RECV-MESSAGE or 2f + 1 GAP-DROP-MESSAGES. If a GAP-RECV-MESSAGE is received, the leader first verifies the enclosed OC following the same procedure as above.

When replica i receives a GAP-DECISION, it verifies the enclosed OC if the decision contains a GAP-RECV-MESSAGE. If the decision contains 2f + 1 GAP-DROP-MESSAGE, the replica verifies that all 2f + 1 messages are from distinct replicas, and their log-slot-num matches the one in the GAP-DECISION. It then broadcasts a \langle GAP-PREPARE, view-id, i, log-slot-num, recv-or-drop\rangle_{\sigma}, where recv-or-drop is a binary value indicating the decision, to all replicas.

Once replica i receives 2f GAP-PREPARES from distinct replicas (possibly including itself) and it has received a validated GAP-DECISION with a matching decision from the leader, it broadcasts a \langle GAP-COMMIT, view-id, log-slot-num, recv-or-drop\rangle_{\sigma} to all replicas. The replica also stores the GAP-DECISION and 2f GAP-PREPARE in its log for the view change protocol.

When replica i receives 2f GAP-COMMITS from different replicas (possibly including itself), it stores the log (if it hasn’t done so) or a NO-OP to the log slot based on the decision, and resuming normal operation if it is blocking on a QUERY-REPLY or a gap agreement decision. It also stores all 2f + 1 GAP-COMMITS in its log. This quorum of GAP-COMMIT will serve as a gap certificate for the state synchronization and view change protocols. In the rare case where the replica has already speculatively executed the request and the decision is a drop, it rolls back the application state to before log-slot-num, and re-executes subsequent requests in the log.

A.2 View Change Protocol

In this section, we provide the detail of the view change protocol, which is omitted in the main paper.

When replica i initiates a view change, it broadcasts a \langle VIEW-CHANGE, view-id, v', epoch-cert, log\rangle_{\sigma} to all other replicas, where view-id is its current view number, v' is the new view, and epoch-cert contains an epoch certificate for each epoch it has started.

When leader replica l of view v' receives 2f valid VIEW-CHANGE messages for view v' from different replicas, it merges the logs in the VIEW-CHANGE messages as follows:

1. It finds the view-id with the largest epoch number e that is supported by a epoch-cert.
2. If the replica has not started epoch e yet, it finds a valid log that has started epoch e. It then copies all requests and NO-OPS before the starting log position of epoch e to its own log.
3. From all valid logs that have started epoch e, it locates the log with the largest seq-num in epoch e. It then copies all requests in epoch e from the log to its own log.
4. From any valid log that has started epoch e, it copies all NO-OPS in epoch e from the log into its own log, possibly overwriting existing requests.

Leader replica l then broadcasts a \langle VIEW-START, v', view-change-msgs\rangle_{\sigma} to other replicas, where view-change-msgs contains the 2f VIEW-CHANGE messages it uses to merge the log and the VIEW-CHANGE message it would have sent for v'.

When replica i receives a VIEW-START message with v' higher than its current view, it checks that view-change-msgs are properly signed by 2f + 1 different replicas, they all contain the same next view number v', and their logs are valid. It then merges its log with logs in view-change-msgs using the same procedure we described above.

If the view change does not involve a epoch switch, replica i can immediately enter the new view. Otherwise, it broadcasts a \langle EPOCH-START, e', log-slot-num\rangle_{\sigma} to all replicas, where e' is the new epoch number, and log-slot-num is the last log index after merging the logs during view change. Once a replica receives 2f + 1 EPOCH-START messages from different replicas with e' and log-slot-num matching its own, it can enter the new view. It also stores these EPOCH-START messages locally as an epoch certificate for future view changes.

A.3 State Synchronization

During normal operations, replicas execute client requests speculatively before they become durable. A speculatively executed request might be overwritten due to the gap agreement or view change protocols, and the replica has to rollback application state and re-executes all subsequent requests in the log. To further reduce the frequency of roll backs and the number of re-executions, we use a periodic synchronization protocol. The goal of the synchronization protocol is to produce a sync-point, where all log entries before and including the sync-point are committed. A committed log entry will never be overwritten or removed, and will be present in the log (at the same position) of all non-faulty replicas in all subsequent views.

After every N entries are added to the log (N is a configurable constant), a replica i broadcasts a \langle SYNC, view-id, log-slot-num, drops\rangle_{\sigma} to all replicas, where log-slot-num is latest log index that is a multiple of N, and drops contains gap certificates for all log slots that have been committed as NO-OP in the current view. Once replica i receives 2f SYNC
messages with the same log-slot-num from different replicas, for each entry in any of the drops that has a valid gap certificate, it writes a NO-OP (possibly overwriting existing request) to the corresponding log position and saves the gap certificate. It then updates its sync-point to log-slot-num.

B Correctness Proof

This section contains complete safety and liveness proofs of the Matrix protocol.

B.1 Safety

The main safety property guaranteed by Matrix is linearizability. In this safety proof, we assume the network primitive AOM provides transferable authentication, ordering, and drop detection, as specified in the main paper.

Theorem 1 (Matrix Safety). Matrix guarantees linearizability of client operations and returned results.

Before proving Theorem 1, we first define a few properties of Matrix replica logs and client REQUESTs.

Definition. A REQUEST is committed in a log slot if it is executed by 2f+1 distinct replicas in that slot with matching view-id and log-hash.

Definition. A REQUEST is successful if the client receives 2f+1 valid REPLYs from different replicas with matching view-id, log-slot-num, log-hash, and result.

It is easy to see that a successful REQUEST implies that the REQUEST is committed.

Definition. A log is stable if it is a prefix of the log of every non-faulty replica in views higher than the current one.

Lemma 1 (Log Stability). Every successful REQUEST was appended onto a stable log at some non-faulty replica, and the resulting log is also stable.

To prove Lemma 1, we first prove the following set of lemmas.

Lemma 2. All non-faulty replicas that begin an epoch begin the epoch with the same log position.

Proof. Prove by induction. In the first epoch, all non-faulty replicas start with log position 0. This proves the base case. Now assume all non-faulty replicas start epoch e with the same log position. To enter the next epoch e', a non-faulty replica needs to receive 2f+1 EPOCH-START messages for e' from distinct replicas with log-slot-num matching its own. Define these 2f+1 EPOCH-STARTS as a epoch certificate. By quorum intersection, no two non-faulty replicas can have epoch certificates with different log-slot-num. Therefore, all non-faulty replicas enter epoch e' with the same log position. This proves the inductive step.

Lemma 3. Within an epoch, if a REQUEST is committed in some log slot l, then no replica can include a NO-OP with a valid gap certificate in slot l in that epoch.

Proof. We prove by contradiction. Assume a replica inserts a NO-OP with a valid gap certificate in slot f. The replica then has received 2f+1 GAP-COMMITS with decision drop for slot l, implying some replica has received 2f distinct GAP-PREPAREs containing a drop decision and a matching GAP-DECISION from the leader. Since a non-faulty replica only sends unique GAP-PREPARE for a log slot within a view, by quorum intersection, there cannot exist 2f distinct GAP-PREPAREs containing a recv decision at slot l. Therefore, drop is the only possible commit decision for the gap agreement protocol. Moreover, a valid GAP-PREPARE containing a drop decision implies that 2f+1 replicas have sent a GAP-DROP-MESSAGE. Out of those 2f+1 replicas, at least f+1 are non-faulty. Since non-faulty replicas block until they receive GAP-COMMIT and the only possible commit outcome is drop, they will not execute REQUEST. By quorum intersection, no 2f+1 replicas can execute REQUEST, so REQUEST cannot be committed. This leads to a contradiction.

Lemma 4. For any two non-faulty replicas in the same epoch, no slot in their logs contains different REQUESTs.

Proof. Within the epoch, non-faulty replicas insert REQUESTs into their logs strictly in the order received from AOM. In the absence of DROP-NOTIFICATIONS, the ordering property of AOM ensures that all non-faulty replicas have identical sequence of REQUESTs in their logs. By Lemma 2, for any epoch, all non-faulty replicas start the epoch with the same log position. Consequently, for any two non-faulty replicas, no log slot within the epoch contains different REQUESTs. A non-faulty replica may also insert a REQUEST r into its log when handling a DROP-NOTIFICATION. To fill the gap caused by the DROP-NOTIFICATION, the replica requires the transfer of r with the corresponding ordering certificate OC (through QUERY-REPLY or GAP-COMMIT). The transferable authentication property of AOM ensures that r is identical to REQUESTs delivered by other non-faulty replicas at the same position in the AOM message sequence. The case is therefore equivalent to the case in which a REQUEST, not a DROP-NOTIFICATION, is received by the replica. A non-faulty replica may also insert a NO-OP into its log during the gap agreement protocol. If the replica inserts NO-OP at the l+i-th log slot where l is the starting log position of the epoch. The replica ignores the corresponding REQUEST by checking the sequence number if it is later delivered by AOM. Consequently, if the replica receives the i+1-th REQUEST in the AOM message sequence, it can only insert the REQUEST at log position l+i+1. By the above argument, the replica’s log from l+i+1 onward will not contain non-matching REQUESTS from other non-faulty replicas. By induction on i, no log slot within the epoch contains different REQUESTS at any two non-faulty replicas.
Lemma 5. Any \texttt{REQUEST} or \texttt{NO-OP} that is committed at a log position in some view will be in the same log slot in all non-faulty replica’s log in all subsequent views.

\textit{Proof.} To enter a new view \(v’\) from the current view \(v\), a non-faulty replica needs to receive a \texttt{VIEW-START} message which contains \(2f+1\) \texttt{VIEW-CHANGE} messages from distinct replicas. There are two cases. Case 1: If a request \(r\) is committed at log position \(i\) in view \(v\), by definition, \(r\) is executed by \(2f+1\) replicas at the same log position in \(v\). Therefore, at least \(f+1\) non-faulty replicas have inserted \(r\) into their log at position \(i\). By quorum intersection, at least one of the \texttt{VIEW-CHANGE} messages contains \(r\) in the log. The log merging rule in the view change protocol ensures that the replica inserts \(r\) into its log at the same log position in view \(v’\). And by Lemma 3, no \texttt{NO-OP} can be committed at the same log position, so the replica will not overwrite the slot with a \texttt{NO-OP} during log merging. Case 2: If a \texttt{NO-OP} is committed at log position \(i\) in view \(v\), at least \(2f+1\) replicas have sent \texttt{GAP-COMMIT} with decision \texttt{DROP-NOTIFICATION} for log slot \(i\). Therefore, at least \(f+1\) non-faulty replicas have stored \(2f\) \texttt{GAP-PREPARE} and the matching \texttt{GAP-DECISION} with decision \texttt{DROP-NOTIFICATION}. By quorum intersection, at least one of the \texttt{VIEW-CHANGE} messages contains the \(2f\) \texttt{GAP-PREPARE} and the \texttt{GAP-DECISION}. The log merging procedure ensures that slot \(i\) is filled with a \texttt{NO-OP} in view \(v’\). \qed

We are now ready to prove the main log stability lemma.

\textit{Proof of Log Stability (Lemma 1).} A successful \texttt{REQUEST} implies that \texttt{REQUEST} is committed at log slot \(l\). By definition of committed requests, at least \(f+1\) non-faulty replicas have matching logs up to \(l\). And since non-faulty replicas insert log entries strictly in log order (blocking before the next log entry is resolved), all log entries before \(l\) are occupied. For any log slot \(l’ \leq l\), if a \texttt{REQUEST} \(r\) is stored in the matching logs, by Lemma 4, no other \texttt{REQUEST} can be inserted into the same slot at any non-faulty replica. And by quorum intersection and our gap agreement protocol, there cannot exist a valid gap certificate for slot \(l’\). Therefore, only \(r\) can be committed at log slot \(l’\) in the view. Otherwise, if a \texttt{NO-OP} is stored in the matching logs, there must exist a valid gap certificate for slot \(l’\). By definition, the \texttt{NO-OP} is committed at \(l’\). Lemma 5 then guarantees that log entry at \(l’\) (either a \texttt{REQUEST} or a \texttt{NO-OP}) in the matching log will be in the same log slot in all non-faulty replica’s log in all subsequent views. Consequently, the successful \texttt{REQUEST} was appended onto a stable log at least \(f+1\) non-faulty replica. Since the successful \texttt{REQUEST} is also committed at log slot \(l\), by Lemma 5, \texttt{REQUEST} will be in slot \(l\) in all non-faulty replica’s log in all subsequent views. The resulting log is thus also stable. \qed

Proof of Matrix Safety (Theorem 1). First, observe that, by definition, a stable log only grows monotonically. Combining this observation with Lemma 1, from the client’s perspective, the behavior of Matrix is indistinguishable from the behavior of a single, correct machine that processes \texttt{REQUEST} sequentially. This implies that any execution of Matrix is equivalent to some serial execution of \texttt{REQUESTS}. Moreover, clients retry sending an operation until a \texttt{REQUEST} containing the operation is successful. Matrix applies standard at-most-once techniques to avoid executing duplicated \texttt{REQUESTS}. Therefore, a Matrix execution is equivalent to some serial execution of unique client operations.

The above argument proves serializability. When a client receives the necessary replies for a successful \texttt{REQUEST} \(r\), by Lemma 1, \(r\) must have already been added to a stable log. For any successful \texttt{REQUEST} \(r’\) issued after this point in real-time, \(r’\) can only be inserted after \(r\) in the stable log. Since non-faulty replicas execute \texttt{REQUESTS} strictly in log order, the operations issued and results returned by Matrix are linearizable. \qed

B.2 Liveness

Due to the well-known FLP result, Matrix can not guarantee progress in a fully asynchronous network. We therefore only prove liveness given some weak synchrony assumptions.

\textbf{Theorem 2 (Liveness).} \texttt{REQUESTS} sent by clients will eventually be successful if there is sufficient amount of time during which

\begin{itemize}
  \item the network the replicas communicate over is fair-lossy,
  \item there is some bound on the relative processing speeds of replicas,
  \item the \(2f+1\) non-faulty replicas stay up,
  \item there is a non-faulty replica that stays up which no non-faulty replica suspects of having failed,
  \item there is a non-faulty AOM sequencer stays up which no non-faulty replica suspects of having failed,
  \item all non-faulty replicas correctly suspect faulty nodes and AOM sequencers,
  \item clients’ \texttt{REQUESTS} are eventually delivered by AOM.
\end{itemize}

\textit{Proof.} Since there exist non-faulty replica and non-faulty AOM sequencer that stay up which no non-faulty replica suspects of having failed, there is a finite number of view changes during the synchrony period. Once the non-faulty replica that stays up has been elected as leader, and the non-faulty sequencer has been configured by the network, no view change with a higher view will start, as \(2f+1\) non-faulty replicas will not send the corresponding \texttt{VIEW-CHANGE} message.

Moreover, any view change that successfully starts will eventually finish. A non-faulty replica that initiates a view change keeps re-broadcasting its \texttt{VIEW-CHANGE} message until the new view starts, or until the view change is supplanted by one with a higher view. Since non-faulty replicas correctly
suspect faulty nodes and AOM sequencers, eventually $2f + 1$ non-faulty replicas will initiate the view change. As all $2f + 1$ non-faulty replicas stay up, the leader for the new view, if it is non-faulty, will eventually receive the necessary $2f + 1$ VIEW-CHANGE messages to start the view. If the leader is faulty, non-faulty replicas will correctly suspect the fact, and start a higher view change which will supplant the current one.

Additionally, once a view starts, eventually all non-faulty replicas will adopt the new view and start processing REQUESTs in the view, as long as the view is not supplanted by a even higher view. If the leader is non-faulty, it will re-broadcast VIEW-START messages until it receives acknowledgement from all replicas. If the leader is faulty, non-faulty replicas will correctly suspect the fact, and start a higher view change which will supplant the current one.

The above arguments imply that eventually, there will be a view which stays active with a non-faulty leader and a non-faulty AOM sequencer. During that view, non-faulty replicas will eventually be able to resolve any DROP-NOTIFICATION from AOM: The replica receiving a DROP-NOTIFICATION will keep resending the QUERY message until receiving a QUERY-REPLY from the leader or enough GAP-COMMITS. If the leader does not have the REQUEST, it will continually broadcast GAP-FIND-MESSAGE to all replicas. Since $2f + 1$ non-faulty replicas stay up, eventually the leader will receive either one GAP-RECV-MESSAGE or $2f + 1$ GAP-DROP-MESSAGES. Once the non-faulty leader starts the binary Byzantine agreement protocol with the decision, by applying the same line of reasoning, eventually non-faulty replicas blocking on the DROP-NOTIFICATION will receive the necessary GAP-COMMITS to resolve the DROP-NOTIFICATION.

Therefore, the system will eventually reach a point where a view stays active with a non-faulty leader and a non-faulty AOM sequencer, and non-faulty replicas only receive REQUESTs from AOM. After that point, every REQUEST delivered by AOM will eventually be successful. Because clients’ REQUESTs eventually will be delivered by AOM and $2f + 1$ non-faulty replicas stay up, clients will eventually receive the necessary REPLYs for REQUESTs they have sent.