

# MassBFT: Fast and Scalable Geo-Distributed Byzantine Fault-Tolerant Consensus

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**Abstract**—Geo-distributed consensus protocols provide high availability and resilience for distributed database services. These protocols group nodes by their data centers to leverage the network topology that spans across multiple data centers, thereby reducing costly cross-datacenter communication. However, they still face performance and scalability challenges due to inefficient log replication mechanisms. 1) These protocols rely on the leader node in each group to perform cross-datacenter log replication, creating a single-node performance bottleneck. 2) Byzantine receivers can behave arbitrarily, forcing the group leader to send multiple log copies during replication to prevent loss, thus causing redundant transmissions. 3) Since all groups must execute these logs in the same order, synchronizations across groups are necessary to maintain consistency when multiple groups are proposing concurrently, which also slow down log replication.

This paper presents MassBFT, a Byzantine fault-tolerant geo-consensus protocol that achieves high performance and scalability. We design an encoded bijective log replication to eliminate the leader bottleneck and reduce the cross-datacenter network consumption. We also propose asynchronous log ordering to eliminate synchronization across groups. Experimental results show that MassBFT is scalable, fault-tolerant, and outperforms state-of-the-art protocols with 5.49-29.96 times higher throughput under YCSB, SmallBank, and TPC-C workloads.

**Index Terms**—Replication, Fault Tolerance, Consensus

## I. INTRODUCTION

Byzantine fault-tolerant (BFT) consensus protocols are designed to provide availability and fault tolerance for distributed database services, including edge computing [1] [2], cross-border cooperation [3] [4], distributed robust graph storage [5], and authenticated query execution [6]. These services typically span multiple regions and data centers, handle large amounts of requests, and serve a widely distributed user base. For example, in edge computing [7] [8], such as healthcare [9], smart cities [10], and industry automation [11], edge servers are distributed across multiple geographic data centers and require mutual trust to maintain consistent database states. Geo-distributed consensus protocols have emerged as a scalable solution to meet the high demands of such applications.

Geo-consensus protocols, such as Steward [12], Blockplane [13], and GeoBFT [14], utilize high-speed local area networks (LAN) within data centers to reduce the costly wide-area network (WAN) communications between data centers. Instead of using all-to-all communication [15], these protocols adopt a hierarchical structure where each data center group elects a leader to handle log replication with other groups. Figure 1a shows an example of log replication, where group  $G_1$  sends a log entry  $e$  to group  $G_2$ . This process involves three steps:

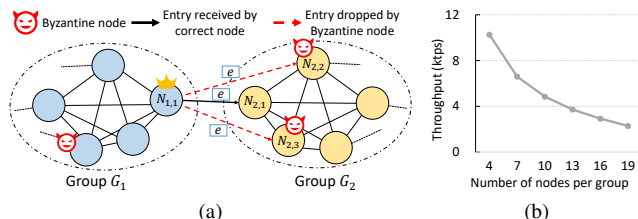


Fig. 1. (a) An example of inter-group log replication. (b) Throughput under different group size settings.

- Batching:** The leader node  $N_{1,1}$  of group  $G_1$  creates an entry  $e$  containing a batch of transactions from local clients.
- Local Replication:** To ensure the integrity of entry  $e$ ,  $N_{1,1}$  broadcasts entry  $e$  within its group via LAN using local BFT consensus (e.g., PBFT [15] [16]).
- Global Replication:** As there are at most  $f$  indistinguishable Byzantine nodes in  $G_2$  that may discard entry  $e$ ,  $N_{1,1}$  sends entry  $e$  to  $f + 1$  nodes in  $G_2$  to ensure that at least one *correct* node (i.e.,  $N_{2,1}$ ) receives it. Upon receipt,  $N_{2,1}$  broadcasts  $e$  locally via LAN to ensure all correct nodes in  $G_2$  receive it. The entries are then executed in a deterministic order across all nodes [17].

Geo-consensus protocols reduce WAN traffic at the expense of increasing LAN traffic. However, they still face significant scalability challenges due to limitations in leader-based log replication and inter-group synchronization, especially when cross-datacenter WAN transmissions become bottlenecks.

First, only the leader node broadcasts entries during global replication, creating a single-node bottleneck due to its limited upstream bandwidth. Although using multiple nodes in group  $G_1$  to transmit an entry can alleviate this bottleneck, ensuring reliable transmission is still challenging. In Figure 1a, when using only  $G_1$ 's leader node  $N_{1,1}$  to send the entry, nodes in  $G_2$  can easily identify any faulty behavior from  $N_{1,1}$ . However, when multiple sending nodes are used, pinpointing the cause of entry loss becomes challenging, as both faulty sending and receiving nodes can tamper with the entry (Section IV-A).

Second, in the broadcast example described above, massive redundant copies of entries are transmitted, resulting in high WAN traffic consumption. In Figure 1a,  $N_{2,2}$  and  $N_{2,3}$  in the receiver group  $G_2$  are faulty.  $N_{1,1}$  must send three copies of entry  $e$  to ensure that  $G_2$  receives at least one correct copy. As the number of nodes in  $G_2$  increases, the number of faulty nodes  $f$  that  $G_2$  can tolerate also increases. Therefore, the number of copies  $N_{1,1}$  sends increases linearly, causing the exhaustion of its upstream bandwidth. Figure 1b shows the scalability issue on GeoBFT [14], which employs the above replication strategy. We deploy 12 to 57 nodes across three

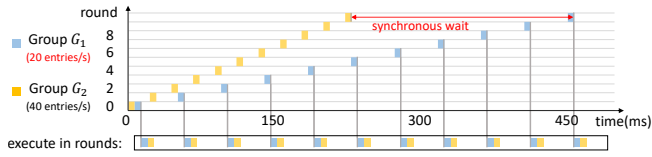


Fig. 2.  $G_2$ 's performance is limited by  $G_1$  due to the round-based ordering.

data centers in our nationwide cluster (detailed setup in Section VI), with each data center hosting 4 to 19 nodes. Each node is equipped with an independent WAN connection limited to 20 Mbps. The result shows a significant decrease in throughput as the number of nodes in each group increases.

Third, synchronizations are required when multiple group leaders propose their entries concurrently. Some protocols [18]–[20], such as Steward [12] and Blockplane [13], utilize Paxos [21] among all group leaders, allowing only one group leader to propose entries at the same time. Other protocols, such as GeoBFT and RCanopus [22], solve this single-leader bottleneck by allowing all group leaders to propose entries concurrently. To order the log entries proposed by different groups consistently, they adopted a predefined strategy that operates in rounds: in each round, every group must propose exactly one entry, and entries proposed in the same round are executed in order by comparing their group IDs.

This *synchronized* round-based ordering presumes the entry replication rates of all groups are fixed and identical. Unfortunately, even when the network between groups is stable, some groups may replicate entries slower than others due to limited WAN bandwidth or skewed workload distribution. As a result, entries proposed by the fast groups with a later execution order may arrive earlier. However, according to the predefined order, these entries can only be executed once all their predecessors have arrived and executed, limiting the throughput of the faster groups and increasing latency. In Figure 2,  $G_1$  proposes 20 entries per second while  $G_2$  proposes 40. In each round, entries from  $G_2$  must wait for those from  $G_1$  before they can be executed, thereby limiting  $G_2$ 's performance.

To address these challenges, we propose MassBFT, a geo-consensus protocol that significantly enhances scalability.

We propose encoded bijective log replication for transmitting entries between groups, enabling all nodes to participate in the process, thus preventing group leaders from becoming single-node bottlenecks. In contrast to previous group sending approaches [14] [23] [24], where each node transmits complete entry copies (Section IV-A), we employ erasure coding [25] to reduce the number of entry copies transmitted between groups. Entries are divided into small *data chunks*, and additional *parity chunks* are encoded to ensure successful recovery in the event of data chunk loss due to node failures. By carefully designing a transfer plan, each node transmits only a few chunks instead of a complete entry copy while ensuring that the receiving group obtains enough chunks to reconstruct the entry, thereby reducing WAN traffic (Section IV-B).

We asynchronously determine the execution order of entries after they finish global replication, enabling entries proposed by faster groups to execute without delay. Each entry is assigned a vector timestamp (VTS) after completing global

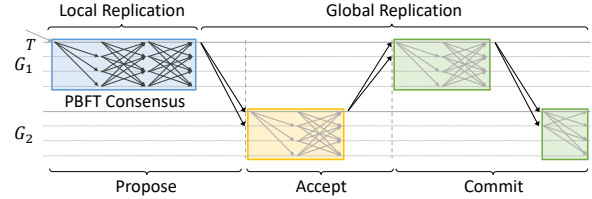


Fig. 3. Log replication of group  $G_1$  in Baseline.

replication, and the *chronological order* of these VTSs dictates their execution order. Entries that are replicated earlier have smaller VTS and can be executed first, without waiting for other entries to finish replication. Replicating VTS is non-blocking and lightweight, adding negligible WAN traffic. Thus, even if a slower group has not yet received an entry, it can still replicate the entry's VTS on schedule (Section V-C). The VTSs are deterministically ordered to establish a consistent execution order among nodes (Section V-D).

In summary, the contributions of this paper include:

- We propose encoded bijective log replication to accelerate entry transmission. Group followers are used to send entries to avoid leader bottlenecks, while erasure coding encodes each entry into fault-tolerant chunks to reduce WAN traffic.
- We propose asynchronous log ordering to eliminate synchronization waiting issues before execution due to discrepancies in entry replication rates among groups.
- We present MassBFT based on the techniques above and conduct comprehensive evaluations. The results show that MassBFT is fast and scalable and outperforms all baselines.

## II. BACKGROUND ON GEO-CONSENSUS

Geo-consensus protocols typically adopt a hierarchical architecture that groups nodes by data centers. They can vary in threat models (crash or Byzantine), structures (two or multi level), and protocols (PBFT [15] [16], Paxos [21], or Raft [30]) for local and global consensus. To clearly outline the workflow of geo-consensus protocols, we introduce Baseline as a generic model and review existing geo-consensus protocols.

### A. Bottleneck Analysis in Log Replication

Baseline requires  $n \geq 3f + 1$  nodes per group, with at most  $f$  out of  $n$  nodes that behave *faulty* (i.e., Byzantine). Since entire groups may crash due to WAN network partitions or data center power outages, Baseline also requires at least  $n_g \geq 2f_g + 1$  groups to maintain liveness, allowing for up to  $f_g$  groups to crash simultaneously. For two groups,  $G_1$  and  $G_2$ , Figure 3 shows the log replication of  $G_1$  in Baseline.  $G_2$ 's process is identical to  $G_1$ 's and omitted. The detailed workflow of Baseline is described below.

**Batching.** As with any multi-master protocols [29] [22] [14], Baseline enables all groups to serve requests simultaneously. To minimize response time and WAN traffic consumption, clients send transaction requests to the nearest group leader based on network latency or physical distance [31] [12]. Once a certain number of requests are received or a timeout occurs, the group leader batches the transactions into a log entry (or simply, an entry). The entry also includes metadata such as *term* and *commitIndex* for global Raft consensus.

TABLE I  
COMPARISON OF GEO-CONSENSUS PROTOCOLS WITH HIERARCHICAL ARCHITECTURE

Protocol	FT	Local consensus	Global consensus	Group failure	Log replication	Multi-master	Global ordering
D-Paxos [18]	CFT	Multi-Paxos	Paxos	$n_g \geq 2f_g + 1$	One-way (leader)	No	-
C-Raft [19]	CFT	Fast Raft	Fast Raft	$n_g \geq 2f_g + 1$	One-way (leader)	No	-
Steward [12]	BFT	PBFT	Paxos	$n_g \geq 2f_g + 1$	One-way (leader)	No	-
Blockplane [13]	BFT	PBFT	Paxos	$n_g \geq 2f_g + 1$	One-way (leader)	No	-
Ziziphus [20]	BFT	PBFT	Paxos	$n_g \geq 2f_g + 1$	One-way (leader)	No	-
ByzCoin [26]	BFT	-	PBFT	$n_g \geq 3f_g + 2$	Multi-level tree	No	-
Kauri [27]	BFT	-	HotStuff	$n_g \geq 3f_g + 1$	Multi-level tree	No	-
ByzCoinX [28]	BFT	-	PBFT	$n_g \geq 3f_g + 1$	Two-level tree	No	-
Canopus [29]	CFT	Raft	-	No	Leaf-only tree	Yes	Synchronous
RCanopus [22]	BFT	BFT+CFT	BFT	$n_g \geq 3f_g + 1$	Leaf-only tree	Yes	Synchronous
GeoBFT [14]	BFT	PBFT	-	No	One-way (leader)	Yes	Synchronous
Baseline	BFT	PBFT	Raft	$n_g \geq 2f_g + 1$	One-way (leader)	Yes	Synchronous
MassBFT (ours)	BFT	PBFT	Raft	$n_g \geq 2f_g + 1$	Encoded bijective	Yes	Asynchronous

**Local Replication.** PBFT is used for achieving local consensus within each group to tolerate Byzantine nodes. In the blue box of Figure 3, the group leader creates a certificate for the entry by running PBFT in three phases (pre-prepare, prepare, and commit) within its group. The certificate protects the entry from tampering by Byzantine nodes during the subsequent global replication. Once PBFT is complete, all correct nodes in the group receive the entry along with the certificate.

**Global Replication.** Baseline adopts Raft for global replication, with each group serving as a logical replica participating in Raft consensus. Since the Raft messages proposed by each group already undergo local consensus and are protected by PBFT certificates, Byzantine nodes cannot tamper with them. Therefore, the global Raft consensus aims to provide fault tolerance when an entire group crashes. As shown in Figure 3, the leader node of  $G_1$  initiates Raft consensus (which proceeds through *propose*, *accept*, and *commit* phases) by broadcasting the entry with the certificate to other groups  $G_2$  through WAN. The leader of  $G_1$  must send the entry to at least  $f + 1$  nodes in  $G_2$  to tolerate at most  $f$  Byzantine nodes in  $G_2$  from dropping it. When the leader node of  $G_1$  does not respond,  $G_2$  can initiate a remote view change protocol [14] [20] to replace it. The correct nodes then forward the entry to all nodes within its group  $G_2$  via LAN. When the leader of  $G_2$  receives the entry, it creates an *accept* message as a receipt. To prevent Byzantine decisions from the leader node, nodes in  $G_2$  reach PBFT consensus on the *accept* message before replying to  $G_1$ . As shown in the yellow box in Figure 3, PBFT consensus skips the prepare phase because nodes in  $G_2$  do not need to agree on the consensus input (i.e., the entry), as it has already been certified by nodes in  $G_1$  (details can be found in Ziziphus [20]). When the leader of  $G_1$  receives *accept* messages from  $f_g + 1$  groups (i.e.,  $G_2$ ), it creates a *commit* message to support these *accepts* and reaches PBFT consensus on *commit* in  $G_1$ . After that, it broadcasts *commit* to other groups and assumes that the entry is replicated. Nodes in  $G_2$  commit the entry after receiving the *commit* message.

Meanwhile, entries proposed by  $G_2$  replicate to  $G_1$  in a similar way, as all groups accept requests concurrently. To order the entries proposed by  $G_1$  and  $G_2$  consistently across all nodes (a consistent execution order is crucial for accurate results), Baseline adopts a round-based synchronous ordering that proceeds in rounds. Every group can propose

exactly one entry per round. After receiving all entries in a round, a node orders these entries by the IDs of the groups that created them and executes these entries in this order.

### B. Comparison of Geo-Consensus Protocols

Table I compares existing geo-consensus protocols with Baseline mentioned above and our protocol, MassBFT.

**Consensus Architecture.** The choice of the local consensus protocol highly depends on their failure models. D-Paxos [18], Canopus [29], and C-Raft [19] only tolerate crash failures because they use crash fault-tolerant (CFT) protocols (i.e., Multi-Paxos [32], Raft, and Fast Raft [19]) for local consensus. Canopus groups nodes within the same rack to optimize the performance of local consensus. C-Raft uses Fast Raft to reduce the number of network round-trips (RTTs) during consensus. In contrast, Steward, Blockplane [13], Ziziphus, and GeoBFT use PBFT for local consensus to tolerate Byzantine failures. RCanopus [22] is similar to that of Canopus. Nodes within the same rack are organized into CFT sub-groups. Each group, consisting of multiple geographically close sub-groups, uses BFT consensus to tolerate Byzantine sub-groups.

The goal of global consensus is to reduce network traffic between groups. There are three types of global consensus protocols: 1) Protocols like D-Paxos and Steward use CFT global consensus, which can tolerate up to  $\lfloor (n_g - 1)/2 \rfloor$  crashed groups  $f_g$  at the same time. 2) Protocols like ByzCoin [26] and RCanopus use BFT global consensus, which can cope with Byzantine groups. Moreover, ByzCoin, Kauri [27], and ByzCoinX [28] do not employ local consensus. The group leaders serve as intermediaries, disseminating the proposal from the PBFT leader to their group members and aggregating their responses back to the PBFT leader. 3) GeoBFT and Canopus replicate logs directly without using global consensus. After achieving local consensus, the entry is directly broadcast to other groups. This approach can significantly reduce latency but at the cost of sacrificing group fault tolerance (i.e., they cannot tolerate Byzantine or crash groups).

**Log Replication.** Most existing protocols [12]–[14] adopt a one-way replication strategy where only group leaders initiate entry transfer. Moreover, they must send entry copies to additional receivers to avoid entry loss during replication. For example, the group leaders in Steward send each entry to all nodes, while GeoBFT reduces WAN overhead at the cost of

additional intra-group communication, where the entry is only sent to  $f + 1$  nodes within each group across all other groups. ByzCoin and Kauri employ multi-level trees to reduce WAN transmission overhead at the expense of latency. Additionally, Kauri employs HotStuff [33] to reduce message complexity and leverages pipelining techniques to enhance throughput. However, HotStuff may experience latency bottlenecks in geo-distributed scenarios. ByzCoinX employs a two-level tree to balance the latency and the inter-group network traffic. Canopus and RCanopus utilize a Leaf-only Tree overlay for entry replication, which supports pipelining but also increases latency due to multiple rounds of entry exchanges.

**Ordering.** Steward and Blockplane limit only one group to propose entries at a time. For example, in Steward, when multiple group leaders propose entries at the same time, only one (or none) can succeed because Paxos can only reach a consensus on a single entry at a time. Therefore, the execution order of entries is exactly the output of the global consensus. D-Paxos reduces contention in Paxos by using a round-robin strategy, where group leaders take turns acting as the global consensus leader to propose entries. However, D-Paxos still allows only one group leader to propose entries at a time.

In contrast, Canopus, RCanopus, and GeoBFT enable all groups to propose entries simultaneously. When multiple groups propose entries at the same time, nodes may receive these entries in any order. To establish a global consistent execution order, these protocols operate in rounds to order these entries. In each round, every group proposes exactly one entry. Execution at a node begins once all entries for the current round have been received. However, this method requires all groups to replicate entries at the same rate.

### III. MASSBFT OVERVIEW

#### A. System Model

MassBFT is deployed on geo-distributed data centers. Nodes within the same group are on the same data center and are connected through a fast data center network (LAN). Each node also has an exclusive public network (WAN) with limited bandwidth. We denote the  $i$ -th group as  $G_i$ , and the  $j$ -th node in the  $i$ -th group  $G_i$  as  $N_{i,j}$ .

**Threat Model.** MassBFT adopts the BFT model, where *faulty* nodes can behave arbitrarily. A strong adversary can control all faulty nodes but cannot break the cryptographic primitives. MassBFT adopts a public-key infrastructure (PKI), where each node has a public-private key pair for signing and verifying messages. We denote the number of faulty nodes in group  $G_i$  as  $f_i$ , and the number of nodes  $n_i$  in group  $G_i$  is at least  $3f_i + 1$ . The entire group is crash-only, and MassBFT requires  $n_g \geq 2f_g + 1$  groups, with up to  $f_g$  groups crashed simultaneously.

**Network Model.** For *liveness*, MassBFT assumes the partial synchrony model [34]. The network is considered to be reliable most of the time but may also experience periods of asynchrony. During each unstable period, a global stabilization time (GST) exists, after which all transmissions between correct nodes arrive within a known bound  $\Delta$ .

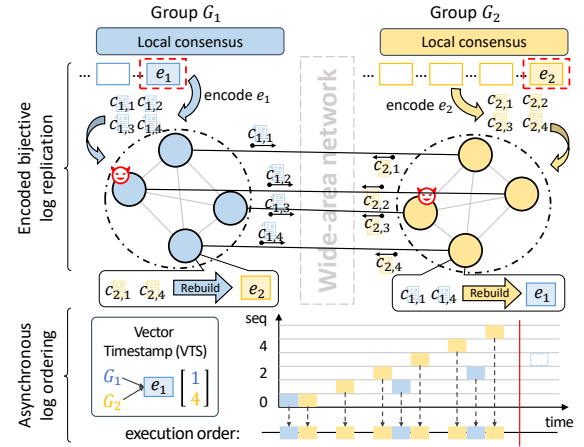


Fig. 4. Overview of MassBFT.

#### B. Protocol Overview

Figure 4 shows the workflow of MassBFT. The local replication is the same as that of Baseline (Section II-A). During global replication, entries are transferred via encoded bijective log replication, followed by asynchronous log ordering to establish a consistent execution order. The key features of MassBFT are described below.

##### Multi-master Replication with Hierarchical Architecture.

In MassBFT, all groups concurrently serve requests and propose entries. First, entries containing clients' transactions undergo local PBFT consensus within the group to mask the behaviors of faulty nodes. Then, the entries are replicated among all groups using Raft to tolerate crashed groups. After Raft consensus, entries generated by the same group are already in order (proven in Section V-D). For entries generated by different groups, MassBFT adopts the asynchronous log ordering to ensure a consistent execution order across nodes, preventing slow groups from slowing down others.

##### Bijective Low-redundancy Log Replication.

MassBFT utilizes all nodes to transmit entries and employs erasure coding to minimize redundancy. We now consider sending an entry between two groups, as broadcasting entries to all groups can be built from it. In Figure 4, after achieving local consensus, every correct node in group  $G_1$  has a copy of entry  $e_1$ . Each node in  $G_1$  then deterministically splits  $e_1$  into data chunks, and encodes additional parity chunks to tolerate chunk loss caused by faulty nodes in both groups. Each chunk is sent only once from one node in  $G_1$  to another in  $G_2$  based on a transfer plan to reduce redundancy. Finally, correct nodes in  $G_2$  exchange their received chunks, ensuring each node has sufficient chunks to rebuild entry  $e_1$ . The replication process of  $e_2$  is similar than that of  $e_1$ . Notably, only the *propose* phase, which handles entry replication, is enhanced with encoded bijective log replication, while the *accept* and *commit* phases remain the same as Baseline.

##### Asynchronous Ordering with Vector Timestamps.

To tolerate slow groups, MassBFT orders entries by their logical vector timestamps (VTS), which are assigned after they finish global replication. To avoid centralized timestamp assignment,



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**Algorithm 1: Transfer Plan Generation**  
 (for each sender-receiver group pair)
 

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**Input:** node with ID  $i$  (in the sender group or receiver group), sender group  $G_1$  of size  $\mathbf{n}_1$ , and receiver group  $G_2$  of size  $\mathbf{n}_2$ .

**Output:** the number of total chunks  $n_{\text{total}}$  and data chunks  $n_{\text{data}}$ . The transfer plan *plan* with a list of tuples  $\langle \text{chunk ID, sender node ID, receiver node ID} \rangle$ .

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1  $n_{\text{total}} \leftarrow \text{LCM}(\mathbf{n}_1, \mathbf{n}_2)$ ; // calculate the least common multiple
2  $nc_1 \leftarrow n_{\text{total}}/\mathbf{n}_1$ ; // number of chunks a node sends in  $G_1$ 
3  $nc_2 \leftarrow n_{\text{total}}/\mathbf{n}_2$ ; // number of chunks a node receives in  $G_2$ 
4  $\mathbf{f}_1 \leftarrow \lfloor (\mathbf{n}_1 - 1)/3 \rfloor$ ;  $\mathbf{f}_2 \leftarrow \lfloor (\mathbf{n}_2 - 1)/3 \rfloor$ ; // faulty nodes
5  $n_{\text{parity}} \leftarrow nc_1 \times \mathbf{f}_1 + nc_2 \times \mathbf{f}_2$ ; // number of chunks lost
6  $n_{\text{data}} \leftarrow n_{\text{total}} - n_{\text{parity}}$ ; // number of chunks delivered
7 if node with ID  $i$  is in the sender group  $G_1$  then
8   for chunk with ID  $c$  from  $nc_1 \times i$  to  $nc_1 \times (i + 1) - 1$  do
9      $j \leftarrow \lfloor c/nc_2 \rfloor$ ; //  $c, i, j$  start from 0
10    Add  $\langle c, i, j \rangle$  to plan; // send chunk  $c$  to node  $j$  in  $G_2$ 
11  else // node with ID  $i$  in the receiver group  $G_2$ 
12    for chunk with ID  $c$  from  $nc_2 \times i$  to  $nc_2 \times (i + 1) - 1$  do
13       $j \leftarrow \lfloor c/nc_1 \rfloor$ ;
14      Add  $\langle c, j, i \rangle$  to plan; // receive  $c$  from node  $j$  in  $G_1$ 
15  return  $n_{\text{total}}$ ,  $n_{\text{data}}$ , and plan;

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**Case Study.** Figure 5b shows an example where a 4-node group  $G_1$  sends an entry to a 7-node group  $G_2$ . The total number of chunks  $n_{\text{total}} = 28$  is determined by the LCM of 4 and 7. Each  $G_1$  node sends 7 chunks, while each  $G_2$  node receives 4 chunks. Considering 1 faulty node in  $G_1$  and 2 faulty nodes in  $G_2$ , the maximum potential chunk loss is  $1 \times 7 + 2 \times 4 = 15$ , matching the number of parity chunks  $n_{\text{parity}}$ . Based on the transfer plan, correct  $G_2$  nodes receive a total of 13 correct chunks (indicated by gray lines) and broadcast these chunks via LAN. Finally, each  $G_2$  node can obtain at least 13 chunks, which is sufficient to rebuild the original entry. In the example above, the encoded bijective approach sends a total of  $n_{\text{total}}/n_{\text{data}} \approx 2.15$  entry copies, which is significantly lower than the bijective-only approach (Figure 5a), which is 4.

### C. Optimistic Entry Rebuild

During chunk transmission, faulty senders may send fake chunks to receivers, while faulty receivers may also broadcast fake chunks within their groups. The fake chunks are indistinguishable from the correct ones, as they do not undergo local consensus and produce a certificate. Since erasure coding requires all input chunks be correct and properly ordered, receivers might need to rebuild the entry multiple times. This repeated process can significantly affect the liveness property.

We adopt an optimistic approach to validate the received chunks, inspired by previous studies [38]–[40]. This approach leverages Merkle trees and Merkle proofs [41] [42] to efficiently verify whether the received chunks are encoded from the same entry. A Merkle tree is an authenticated data structure used to verify a large data set (i.e., a set of chunks). Each leaf node of the Merkle tree contains the hash of an individual data block (i.e., a chunk), and each internal node contains the hash of its two child nodes. These data blocks can be represented by a single Merkle root, which is the root hash

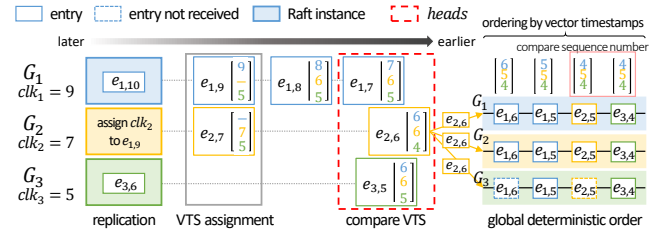


Fig. 6. The workflow of asynchronous log ordering.

obtained by recursively hashing the child nodes. A Merkle proof can efficiently prove whether a data block is included in the data set. The proof consists of node hashes on the path from the leaf node to the Merkle root.

After encoding an entry into chunks, each sender node constructs a Merkle tree from these chunks and sends the entry's PBFT certificate, along with the chunks and their Merkle proofs, to the corresponding receivers. The receivers group these chunks by their Merkle roots into buckets for classification. Tempered chunks must not be grouped into the same bucket as the correct ones. This is because chunks sharing the same Merkle root are encoded from the same entry. When a chunk is tempered, its Merkle root must change.

When the number of chunks in a bucket reaches the threshold (i.e.,  $n_{\text{data}}$ ), the receiver tries to rebuild the entry and validates it with its certificate. If the validation fails, indicating that all chunks in the bucket are fake (as they share the same Merkle root), the receiver logs the IDs of these chunks (inferred from their Merkle proofs). It no longer accepts any further chunks with these IDs to prevent denial-of-service (DoS) attacks. If the entry passes validation, the receiver accepts it and proceeds with the Raft consensus process.

The *liveness* and *safety* properties are discussed as follows:

**Liveness.** The encoded bijective log replication ensures the liveness of inter-group replication [28] [6]. When the network is temporarily unstable and entries are delayed or lost, if the upper protocol provides liveness (e.g., using Raft as global consensus), no additional measures are required. Otherwise (e.g., in GeoBFT), nodes can apply a reliable transmission protocol like TCP to ensure eventual delivery.

**Safety.** Although faulty nodes may tamper with some chunks, as discussed in Section IV-C, the rebuilt erroneous entries can be identified by verifying their PBFT certificates.

## V. ASYNCHRONOUS LOG ORDERING

In MassBFT, the number of nodes in each group may differ due to nodes joining or leaving [2]. Since the entry sending rate of a group is related to its size (verified in Section VI-D) and workload, different groups may propose entries at varying and changing rates. The round-based ordering failed to obtain a flexible execution order, thereby limiting throughput to the pace of the slowest group. Therefore, it is necessary to establish a new total order to execute entries more efficiently.

### A. Protocol Design

Similar to other hierarchical protocols [12] [13], each group in MassBFT participates in the global Raft consensus as a logical replica. As MassBFT adopts a multi-master architecture,

there are a total of  $n_g$  global Raft instances running in parallel. A group  $G_i$  serves as the leader of the  $i$ -th Raft instance (denoted as  $G_i$ 's Raft instance) and participates in all other Raft instances as followers. We denote the entry generated by group  $G_i$  with local sequence number  $m$  as  $e_{i,m}$ .

**Entry Replication.** An entry  $e$  first achieves PBFT consensus within the sending group  $G_i$ . Then,  $G_i$  sends  $e$  to all other groups using the encoded bijective replication. Concurrently,  $G_i$  initiates a global Raft consensus through its own Raft instance, where it acts as the leader. The message includes the entry digest and the corresponding PBFT certificate.

**Vector Timestamp Assignment.** As shown in Figure 6, each group  $G_i$  maintains an independent local logical clock,  $clk_i$ , to timestamp entries. Each entry is assigned a *vector timestamp* (VTS) during the VTS assignment. The global order of entries is determined by comparing their VTSs (Section V-D).

A VTS consists of a set of timestamps. When an entry  $e$  achieves global Raft consensus, each group  $G_i$  assigns the time from its local clock  $clk_i$  to the  $i$ -th element of  $e$ 's VTS, denoted as  $e.VTS[i]$ . This timestamp is then replicated across  $G_i$ 's Raft instance for consistency by piggybacking it onto the Raft consensus during entry replication. Additionally, to ensure liveness, a group increments its local clock once the entry it has proposed completes the global Raft consensus.

For example, Figure 7a shows that after entry  $e_{1,10}$  achieves Raft consensus, the proposer group  $G_1$  increases its local clock  $clk_1$  from 9 to 10. After receiving entry  $e_{1,10}$  from  $G_1$ 's Raft instance, the three groups,  $G_1$ ,  $G_2$ , and  $G_3$  assign their local clock times, 10, 7, and 5 to  $e_{1,10}$ . The groups then exchange their timestamps through their respective Raft instances. Finally, every group possesses a copy of entry  $e_{1,10}$  and its VTS  $\langle 10, 7, 5 \rangle$ .

We overlap an entry's replication and VTS assignment to reduce latency (Section V-B). We also enhance this process to prevent faulty (or slow) groups and nodes from delaying or blocking the assignment of VTSs (Section V-C).

**Ordering by Vector Timestamp.** Each entry is assigned a complete VTS, allowing us to establish their global order by comparing their VTSs (Lemma V.4). However, since entries with VTSs arrive out of order in a streaming manner, additional measures are needed to predict the next entry to execute.

Entries proposed by the same group  $G_i$  (e.g.,  $e_{1,7}$ ,  $e_{1,8}$ , and  $e_{1,9}$  by  $G_1$  in Figure 6) are ordered by their local sequence numbers, since their VTSs increase monotonically (Lemma V.5). Therefore, among the unexecuted entries proposed by group  $G_i$ , entry  $head_i$  with the smallest local sequence number (e.g.,  $e_{1,7}$ ) must precede than the others (e.g.,  $e_{1,8}$  and  $e_{1,9}$ ). As a result, the next entry to execute can be determined by finding the most preceding entry in  $heads = \langle head_1, \dots, head_{n_g} \rangle$  (e.g.,  $e_{1,7}$ ,  $e_{2,6}$ , and  $e_{3,5}$ ). Therefore, a node can recursively select the entry with the smallest VTS in  $heads$  for execution. By default, the VTSs of all entries in  $heads$  must be assigned completely to determine the next entry to execute. To minimize the blocked waiting caused by assigning VTS, a node can infer

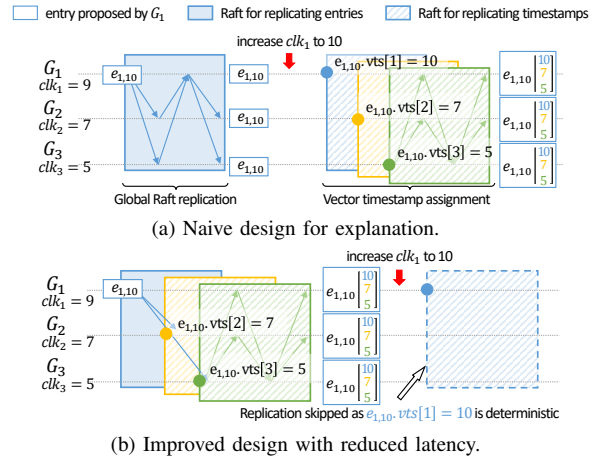


Fig. 7. An illustrative example of entry replication and vector timestamp assignment of entry  $e_{1,10}$ .

the possible VTSs of entries in  $heads$  before they are proposed or before their VTS assignment is finished (Section V-D).

Moreover, concurrent replication may result in entries from different groups having identical VTSs ( $e_{2,5}$  and  $e_{3,4}$ ). A consistent total order is still achievable by comparing their local sequence numbers  $seq$  and group IDs  $gid$  (Section V-D).

### B. Overlapped Vector Timestamp Assignment

As shown in Figure 7a, the VTS assignment discussed above requires two consecutive rounds of Raft consensus, one for replicating the entry and the other for replicating the timestamps assigned to it. The two phases perform serially, causing high consensus latency (about 3 RTTs).

As shown in Figure 7b, it is possible to reduce latency to 2 RTTs while maintaining correctness by overlapping these two phases: **1)** Group  $G_i$  skips assigning the time from its local clock  $clk_i$  to entries  $e_{i,n}$  proposed by itself, because the assignment  $e_{i,n}.vts[i] = n$  is deterministic. Once  $e_{i,n}$  has achieved Raft consensus,  $G_i$  advances its local clock  $clk_i$  to  $n$ . **2)** Every other group  $G_j$  directly assigns the time from its local clock  $clk_j$  to  $e_{i,n}$  on receiving  $e_{i,n}$  via Raft's *propose* message from  $G_i$ , instead of waiting for  $e_{i,n}$  to achieve Raft consensus. However, overlapping the replication phase and the VTS assignment phase of entry  $e$  potentially leads to an exception where some of  $e$ 's VTS assignment succeeds but  $e$ 's replication fails. In this scenario, the correct groups cannot receive  $e$  and therefore cannot execute entries after it, creating blocking issues. However, this exception cannot occur, as proven by the following lemma:

**Lemma V.1. (atomicity)** *If entry  $e$  completes replication, all of its timestamps  $e.vts[j]$  will eventually be assigned and replicated. Otherwise, none of its timestamps  $e.vts[j]$  will complete replication.*

*Proof.* When the timestamp  $e_{i,n}.vts[j]$  from group  $G_j$  is received by group  $G_k$  in a *propose* message,  $G_k$  will respond with an *accept* message if it has received the entry  $e_{i,n}$ . Otherwise,  $G_k$  will not reply until receiving the actual entry  $e_{i,n}$  from group  $G_i$ , or it can request the entry  $e_{i,n}$  from  $G_j$  if group  $G_i$  crashes. Thus, if the timestamp  $e_{i,n}.vts[j]$  completes

replication, at least  $f_g + 1$  groups must have received  $e_{i,n}$  and replied *accept* to  $G_j$ , meaning  $e_{i,n}$  is globally replicated.  $\square$

### C. Handling Faulty Groups and Nodes

**Slow Sender Groups.** We define a slow sender group as a group that proposes and replicates messages at a slower rate than other groups. Although a slow group may take a longer time (tens of milliseconds, depending on the message size and its bandwidth) to send each message, when the network is reliable (due to partial synchrony network assumption), all messages sent by the group can achieve consensus in time.

**Lemma V.2. (liveness)** *After receiving entry  $e_{i,m}$ , group  $G_j$  can assign the current time from its clock  $clk_j$  to  $e_{i,m}.vts[j]$  and finish replicating  $e_{i,m}.vts[j]$  within a short, finite time  $t$ .*

*Proof.* When group  $G_j$  assigns the time from its clock  $clk_j$  to entry  $e_{i,m}$ , another entry  $e_{j,n}$  proposed by  $G_j$  itself may still be replicating. Therefore, the timestamp  $e_{i,m}.vts[j]$  must wait for the replication of  $e_{j,n}$  to complete because they use the same consensus instance. However, the replication of  $e_{i,m}.vts[j]$  cannot be blocked by  $e_{j,n}$  for a long time, because entry  $e_{j,n}$  proposed by  $G_j$  can reach consensus in a short and finite time  $t$  (as previously discussed). We also utilize the pipelining technique to accelerate this process.  $\square$

**Slow Receiver Groups.** A slow receiver group  $G_j$  receives entries  $\langle e_{i,1}, \dots, e_{i,n} \rangle$  slower than the proposing group  $G_i$  sends them. Since these entries must be received in order,  $G_j$  cannot receive and assign its local clock time to the latest entry  $e_{i,n}$  in time. Recall that during Raft consensus, a follower group  $G_k$  replies with an *accept* message to the leader group  $G_i$  when receiving entry  $e_{i,n}$  from  $G_i$  in a *propose* message. To notify the slow group  $G_j$  that  $e_{i,n}$  has completed replication,  $G_k$  also broadcasts the *accept* message directly to other groups (i.e.,  $G_j$ ) without using Raft consensus. The slow group  $G_j$  assigns its local clock time to  $e_{i,n}.vts[j]$  on receiving *accept* messages from  $f_g + 1$  groups (via direct broadcast). This approach avoids slowing down entry ordering of other groups.

**Crashed Groups.** When a group  $G_i$  crashes, its Raft instance (i.e., the Raft instance led by  $G_i$ ) will elect a new leader  $G_j$ . Since  $G_j$  has received the latest log of  $G_i$ 's Raft instance, it can determine  $G_i$ 's clock time  $clk_i = n$  from the local sequence number of the latest entry  $e_{i,n}$  proposed by  $G_i$ .  $G_j$  then represents  $G_i$  by assigning timestamps to entries with  $G_i$ 's clock time (i.e.,  $e.vts[i] = n$ ) and replicates  $e.vts[i]$  using  $G_i$ 's Raft instance. Since  $G_i$  has crashed and cannot propose entries,  $clk_i$  will not increase. When  $G_i$  recovers later,  $G_j$  transfers the leadership of  $G_i$ 's Raft instance back to  $G_i$ , and  $G_i$  can serve requests normally.

**Byzantine Nodes.** Faulty followers are masked via local PBFT consensus, and a faulty leader can be detected by verifying its proposal. For example, if the leader node proposes a stale timestamp, since its local clock must be monotonically non-decreasing, the proposal cannot be validated by local follower nodes and obtain a valid certificate signed by  $2f + 1$  nodes in its group (detailed in Section II-A).

The leader node of a group may exhibit Byzantine behavior by intentionally delaying the sending of Raft messages, while ensuring the delays are shorter than the Raft timeout threshold, reducing consensus performance. To prevent this behavior, we can allow additional  $f$  group followers to send the messages (i.e., message flooding [43] [44]), as the message has reached local consensus and each node in this group has a copy of it.

**Theorem V.3.** *After the replication of entry  $e$  is completed, a complete VTS (with each element been properly set) will be assigned to the entry  $e$  within a finite time  $t$ .*

*Proof.* Lemma V.1 ensures that all replicated entries will be assigned complete VTSs. Lemma V.2 and the failure-handling mechanisms guarantee that these VTSs will complete replication quickly, within a finite time  $t$ .  $\square$

### D. Deterministic Ordering

Differing from vector clocks [45], which are used to determine causal orders, the comparison of VTSs is conducted *element-wise* and terminates once a discrepancy between elements is found. This method is similar to the lexicographical order used in dictionaries. For entries with identical VTSs, we further order them by their local sequence numbers *seq* and group IDs *gid*. For example, in Figure 6, entry  $e_{2,6}$  with VTS  $\langle 6, 6, 4 \rangle$  is ordered before  $e_{3,5}$  with VTS  $\langle 6, 6, 5 \rangle$ , because  $e_{2,6}.vts[1] = e_{3,5}.vts[1] = 6$  and  $e_{2,6}.vts[2] = e_{3,5}.vts[2] = 6$  but  $e_{2,6}.vts[3] = 4$  is smaller than  $e_{3,5}.vts[3] = 5$ .

**Lemma V.4. (total order)** *Let  $e_1$  and  $e_2$  be two entries. We define a strict total order ' $\prec$ ' on the set of all entries such that  $e_1 \prec e_2$  if and only if one of the following conditions holds:*

- $e_1.vts < e_2.vts$
- $e_1.vts = e_2.vts$  and  $e_1.seq < e_2.seq$
- $e_1.vts = e_2.vts$ ,  $e_1.seq = e_2.seq$ , and  $e_1.gid < e_2.gid$

*Proof.* Each group proposes entries in a sequential manner. Therefore, entries proposed by the same group are uniquely distinguished by their respective local sequence numbers *seq*. Since each group proposes a sequence of entries, even if two entries have the same VTS (e.g.,  $e_{3,4}$  and  $e_{2,5}$  in Figure 6), the combination of *seq* and *gid* uniquely identifies and orders any two entries with the same VTS.  $\square$

In real-world scenarios, entries often arrive out-of-order, and different groups also receive these entries in a distinct order. Consequently, to maintain a consistent execution order, before executing an entry  $e$ , it is crucial to ensure that all entries that are ordered before  $e$  have been received and executed.

**Lemma V.5. (monotonicity)** *Let  $e_{i,m}$  and  $e_{i,n}$  be entries proposed by group  $G_i$ . If their local sequence numbers satisfy  $m < n$ , then it must follow that  $e_{i,m} \prec e_{i,n}$ .*

*Proof.* Each group  $G_j$  receives  $e_{i,m}$  and  $e_{i,n}$  and assigns its local clock time to them sequentially in order, ensuring that  $e_{i,m}.vts[j] \leq e_{i,n}.vts[j]$  for any group  $G_j$ . Since  $e_{i,m}.vts \leq e_{i,n}.vts$  and  $e_{i,m}.seq < e_{i,n}.seq$ , according to Lemma V.4, we have  $e_{i,m} \prec e_{i,n}$ .  $\square$

Lemma V.5 can be used to identify the next entry to be executed that is proposed by group  $G_i$ , denoted as *head<sub>i</sub>*.

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**Algorithm 2: Deterministic Ordering by VTS**

---

**Input:** timestamp  $ts_i$  from group  $G_i$ 's clock  $clk_i$  to entry  $e$ ;  
**Output:** sequential execution of deterministically ordered entries via the `Execute` function.

```
1 heads: heads[i] stores the unexecuted entry proposed by
  group  $G_i$  with the smallest local sequence number.
2  $e.set[i]$ : whether  $e.vts[i]$  is inferred (false) or is set (true).
3 Func OnReceiving ( $ts_i, gid, seq$ ): // receive  $e.vts[i]$ 
4  $e \leftarrow \text{GetEntry}(gid, seq)$ ; // find  $e$  by  $gid$  and  $seq$ 
5  $e.vts[i] \leftarrow ts_i$ ;  $e.set[i] \leftarrow true$ ; // set  $e.vts[i]$ 
6 for entry head in heads do // infer head.vts[i] if not set
7   if head.set[i] = false then head.vts[i]  $\leftarrow ts_i$ ;
8   while pre  $\leftarrow \text{GlobalMinimum}()$  and pre is not null do
9     Execute (pre);
10  nxt  $\leftarrow \text{GetEntry}(pre.gid, pre.seq + 1)$ ;
11  heads[pre.gid]  $\leftarrow$  nxt; // replacing entry pre with nxt
12  nxt.vts[nxt.gid]  $\leftarrow$  nxt.seq; nxt.set[nxt.gid]  $\leftarrow true$ ;
13  for  $j$  from 0 to  $n_g - 1$  do // infer nxt.vts element-wise
14    if nxt.set[j] = false then // based on pre.vts
15      nxt.vts[j]  $\leftarrow$  pre.vts[j];
16 Func GlobalMinimum ():
17   for entry  $e_1$  in heads do
18     if (for any other  $e_2$  in heads,  $\text{Prec}(e_1, e_2)$  is true) then
19       return  $e_1$ ; //  $e_1$  is the most ' $\prec$ ' unexecuted entry
20   return null; // other unexecuted entry  $e_2$  may ' $\prec$ ' than  $e_1$ 
21 Func Prec ( $e_1, e_2$ ): // return true if  $e_1$  must ' $\prec$ '  $e_2$ 
22 for  $j$  from 0 to  $n_g - 1$  do // compare VTSs element-wise
23   if  $e_1.set[j] = true$  then
24     if  $e_1.vts[j] < e_2.vts[j]$  then // the increase of  $e_2.vts[j]$ 
25       return true; // won't affect correctness
26     if  $e_2.set[j] = true$  and  $e_1.vts[j] = e_2.vts[j]$  then
27       continue; // their  $vts[j]$  are identical
28   return false; //  $e_2 \prec e_1$  or cannot be compared for now
29 if  $e_1.seq \neq e_2.seq$  then return  $e_1.seq < e_2.seq$ ;
30 return  $e_1.gid < e_2.gid$ ; // compare  $vts, seq$ , and  $gid$ 
```

---

For instance, after executing  $e_{i,1}$ , the next entry to be executed proposed by group  $G_i$  is  $e_{i,2}$ . Therefore, by comparing these next-to-execute entries proposed by each group (i.e.,  $heads = \langle head_1, \dots, head_{n_g} \rangle$ ), one can determine the globally next entry to be executed.

Any two entries  $e_1$  and  $e_2$  in  $heads$  can be ordered by Lemma V.4. To reduce latency, we can establish their order before fully receiving their VTSs. This is achieved by inferring the lower bound for each element in  $e_1.vts$  that has not yet been received, because each group  $G_i$  assigns its local clock time to the entries it receives in a non-decreasing order. Since timestamps are replicated via  $G_i$ 's Raft instance, if another group  $G_j$  receives a timestamp  $e_1.vts[i]$ , any subsequent timestamp  $e_2.vts[i]$  it receives cannot be less than  $e_1.vts[i]$ .

Algorithm 2 shows the pseudo-code of the deterministic ordering. When timestamp  $e.vts[j]$  is inferred, the corresponding bit  $e.set[j]$  is set to *false*. We use `Prec` to determine the order of entries. An entry  $e_1$  orders before an entry  $e_2$  ( $e_1 \prec e_2$ ) if and only if `Prec` returns *true*. `Prec` orders entries based on the following rules:

- When  $e_1.vts[j]$  is inferred,  $e_1$  may not order before  $e_2$ , as the actual  $e_1.vts[j]$  may be greater than  $e_2.vts[j]$ .
- Otherwise, when  $e_1.vts[j]$  is set and  $e_1.vts[j] < e_2.vts[j]$ , the real value of  $e_2.vts[j]$  must be greater than  $e_1.vts[j]$

regardless of whether  $e_2.vts[j]$  is inferred. As a result,  $e_1$  orders before  $e_2$  (lines 23-25).

- We can proceed to compare the next element only when  $e_1.vts[j]$  and  $e_2.vts[j]$  are equal and both have been set. Otherwise, if  $e_2.vts[j]$  is inferred, its value may be greater than  $e_1.vts[j]$ , causing  $e_2$  orders before  $e_1$  (lines 26-28).
- If the VTSs of  $e_1$  and  $e_2$  are identical and both have been set, we use the round-based approach to determine their order by comparing their group IDs  $gid$  and sequence numbers  $seq$  (lines 29-30).

The rest of the algorithm is explained as follows: When receiving a timestamp  $e.vts[i]$  of entry  $e$ , we first update  $e$ 's VTS (lines 4-5) and then infer the  $i$ -th VTS of all entries in  $heads$  (lines 6-7). Next, we recursively find the global next entry to be executed  $pre$  by comparing the entries in  $heads$  using `Prec` (lines 16-20). If such an entry  $pre$  exists, it is executed and replaced by its successor entry  $nxt$  (lines 9-11). If any elements in the VTS of  $nxt$  have not yet been set, they can be inferred from the VTS of  $pre$  (lines 13-15).

**Theorem V.6.** (MassBFT is a consensus protocol) *When the network is partially synchronous, with no more than  $f_g$  groups crashed and no more than  $f$  faulty nodes per group, MassBFT satisfies the following properties:*

**Agreement** MassBFT ensures that all correct nodes execute the same entries in the same order.

**Termination** MassBFT ensures that each correct node eventually completes the execution of entries.

*Proof.* Theorem V.3 ensures entries and their VTSs are sent atomically to all nodes. With identical VTSs as input, Lemma V.4 ensures the order of these entries is deterministic (agreement). We can infer the VTS lower bound of the next entry to be executed (Algorithm 2). As long as at least one group proposes entries, entry  $e$ 's VTS will be ' $\prec$ ' than the VTSs of other unexecuted entries, allowing  $e$  to be executed eventually (termination).  $\square$

## VI. EVALUATION

**Implementation.** We implement a permissioned blockchain prototype using MassBFT for consensus with ~30,000 lines of C++ code. We use BFT-SMART [46] and braft [47] for local and global consensus. Since the C++ erasure code library `liberasurecode` [48] supports encoding up to a maximum of 64 chunks, which does not meet our scalability requirements, our implementation employs a high-performance Go erasure coding library [49]. We employ ED25519 to generate message digests and SHA256 to ensure data integrity. We also leverage pipelining and batching [14] [50] [51] to enhance performance.

Each group concurrently accepts local client transactions and generates a subchain of blocks. These blocks are then synchronized across groups using MassBFT to create a single, globally ordered, ledger. Since we measure end-to-end performance, to prevent non-consensus-related components from becoming bottlenecks, we employ Aria deterministic concurrency control [17] to accelerate transaction execution and use in-memory hash tables to store database states. Finally, clients receive the results of their transactions.

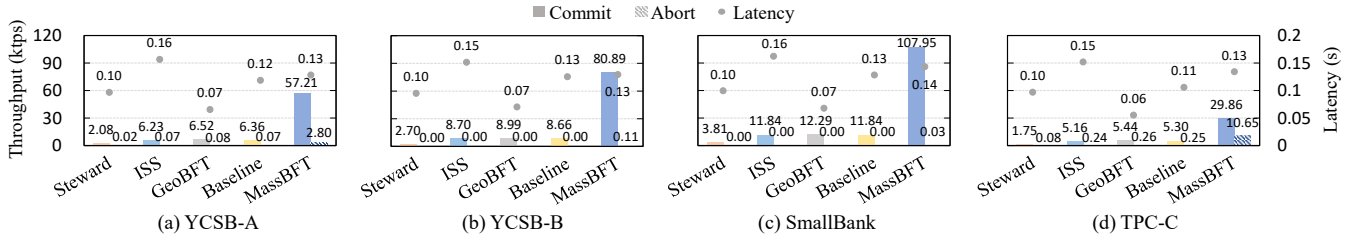


Fig. 8. Performance comparison on the nationwide cluster.

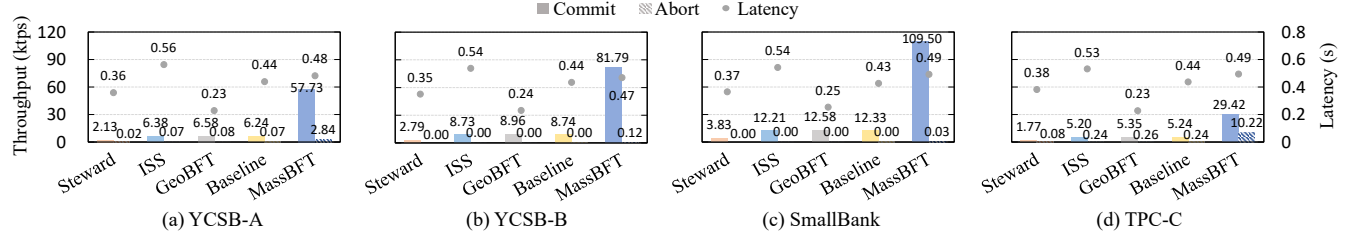


Fig. 9. Performance comparison on the worldwide cluster.

TABLE II  
THE KEY FEATURES OF COMPETITOR SYSTEMS

System	Multi-master	Replication	Consensus	Ordering	Coding
Steward [12]	N	One-way	Raft	-	Entire block
ISS [53]	Y	One-way	Raft+Epoch	Sync.	Entire block
GeoBFT [14]	Y	One-way	Broadcast	Sync.	Entire block
Baseline	Y	One-way	Raft	Sync.	Entire block
MassBFT (ours)	Y	<b>Bijective</b>	Raft	<b>Async.</b>	<b>Erasure-coded</b>

**Physical Environment.** We deploy two geo-distributed clusters on Aliyun [52]. Each cluster consists of 3 groups, with 7 nodes in each group. For the *nationwide cluster*, the groups are located in Zhangjiakou (North China), Chengdu (West China), and Hangzhou (East China). The RTTs between any two groups range from 26.7 ms to 43.4 ms. For the *worldwide cluster*, the groups are located in Hong Kong (Asia), London (Europe), and Silicon Valley (America). The RTTs between any two groups range from 156 ms to 206 ms. Each node (an ecs.c6.2xlarge instance) is equipped with an 8-core CPU, 16 GB of memory, and runs Ubuntu 22.04 LTS. The nodes have independent network interfaces with a WAN bandwidth of 20 Mbps, while nodes within the same data center are connected through LAN with a bandwidth of 2.5 Gbps.

**Competitors.** For a fair comparison, we implement Steward [12], GeoBFT [14], ISS [53], and Baseline (Section II-A) under the same codebase with MassBFT. The batching timeout is fixed at 20ms for all competitors. A comparison of the key features is shown in Table II. Our implementation of Steward uses PBFT for local consensus and Raft for global consensus. GeoBFT uses PBFT for local consensus, and directly broadcasts entries to other groups. Since the Sequenced Broadcast (SB) layer of ISS is pluggable, to ensure a fair comparison, we make ISS a hierarchical consensus protocol by using Steward as its SB layer. The epoch length of ISS is 0.1s. Baseline employs Raft for global consensus, while other configuration parameters remain the same as in GeoBFT. We apply the optimization from GeoBFT to all other protocols. In this case, the group leader only sends the entry to  $f_g + 1$  nodes within each group instead of broadcasting to all nodes. Notably, all other protocols except Steward adopt the multi-master architecture with round-based synchronous ordering.

**Workload.** We use YCSB [54], SmallBank [55], and TPC-C [56] as our experimental workloads. In all workloads, all transactions from clients go through consensus. The YCSB workload is used to evaluate the performance under key-value workloads, using a single table with 10 columns and 1,000,000 rows, with each column having a size of 100 bytes. YCSB transactions follow a Zipf distribution with a skew factor of 0.99. We select YCSB-A (50% read and 50% write) and YCSB-B (95% read and 5% write) workloads. SmallBank workload is used to simulate bank transfer operations. We initialize 1,000,000 accounts, and the access pattern follows a uniform distribution. TPC-C workload is used to simulate order processing operations. We implement a subset of the TPC-C workload that comprises 50% NewOrder and 50% Payment transactions. We initialize 128 warehouses and execute on nodes with 64 GB of memory to accommodate these warehouses. The average transaction sizes of YCSB-A, YCSB-B, SmallBank, and TPC-C are 201B, 150B, 108B, and 232B, respectively. By default, experiments are conducted three times on the nationwide cluster under the YCSB-A workload.

#### A. Overall Performance

Figure 8 and Figure 9 present the performance results on the nationwide and worldwide clusters. MassBFT achieves the highest throughput under all workloads because its global replication strategy is more efficient than the others. ISS, GeoBFT, and Baseline achieve lower throughput due to the single-node bottleneck caused by replicating with group leaders. Steward has the lowest throughput because it allows only one group leader to propose entries at a time.

In Figure 8d, the throughput of MassBFT under TPC-C only increases by 5.64 $\times$  compared to Baseline for two main reasons. First, the transaction signature verification during local PBFT consensus consumes most computing resources, leaving insufficient resources for executing transactions, thus becoming a bottleneck in transaction processing. Second, MassBFT experiences a higher abort rate due to transaction conflicts. This is because the Payment transaction of TPC-C requires accessing hotspot data. As the batching timeout is

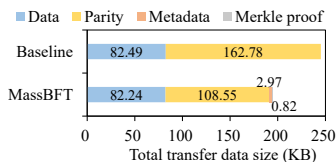


Fig. 10. WAN traffic consumption (YCSB-A).

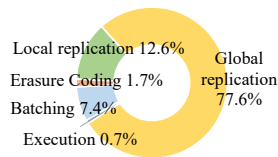


Fig. 11. Latency breakdown (YCSB-A).

fixed, MassBFT processes far more transactions per batch than Baseline (270 vs. 37), leading to a higher abort rate.

In Figure 8a, the latency of MassBFT is slightly higher than that of Baseline (i.e., 128 ms vs. 119 ms). This is because the VTS assignment (Figure 7b) introduces an extra 0.5 RTT. GeoBFT achieves the lowest latency (i.e., 68 ms) because it directly broadcasts entries to all groups without using Raft consensus during global replication, incurring only 0.5 RTT. Baseline has a higher latency than Steward due to round-based synchronization between groups. ISS has the highest latency due to its epoch-based strategy for preventing request duplication, which disrupts the Raft consensus pipeline.

Figure 9 presents the results from the worldwide cluster. All protocols show similar throughput as in nationwide experiments since pipelining effectively hides consensus latency. In Baseline and GeoBFT, the increased distance between nodes makes it difficult to keep groups synchronized, causing additional latency. Although ISS employs per-epoch synchronization to mitigate this issue, the short 0.1s epoch length results in frequent synchronizations, significantly impairing performance. To enhance throughput, we have extended its epoch length to 0.5s. The latency of MassBFT and Steward increases mainly due to the overhead of Raft consensus.

### B. Replication Overhead Analysis

To show the replication efficiency of MassBFT, we measured the network traffic consumption when sending an entry to a remote group. As replication in MassBFT is bijective, the network overhead is the sum of the messages sent by all nodes via WAN. To ensure a fair comparison, we fixed the batch size rather than the batch timeout in this experiment, ensuring that the entry sizes for both MassBFT and Baseline are the same.

As shown in Figure 10, MassBFT consumes less WAN traffic than Baseline when replicating an entry, thanks to the utilization of erasure coding. Although MassBFT must send Merkle proofs for validating the chunks and metadata such as the PBFT certificate for validating the rebuild entry (Section IV-C), these additional costs are negligible.

Figure 11 shows a latency breakdown for MassBFT. Entry encoding and entry rebuild introduce approximately 2.3 ms of latency overhead, which is considered negligible. Most of the overhead comes from global replication, primarily due to high cross-datacenter latency. Additionally, the overhead of local consensus is also significant because nodes must verify the signatures of all transactions to prevent Byzantine faults, which is not required by global Raft consensus.

### C. Results under Heterogeneous Resource

**Groups with different sizes.** In the nationwide cluster, we configure group  $G_1$  with 4 nodes, while group  $G_2$  and group

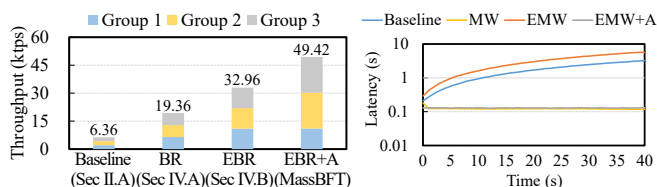


Fig. 12. Results under different sized groups (YCSB-A).

$G_3$  have 7 nodes. We also introduce BR that only applies bijective replication (Section IV-A), EBR that applies encoded bijective replication (Section IV-B), and MassBFT with asynchronous log ordering EBR+A. Figure 12 shows the throughput breakdown and latency results.

Compared to Baseline, BR achieves higher throughput due to decentralized replication. However, its lack of scalability (latency is stable because groups with different sizes have the same throughput) and impracticality in real-world scenarios (Section IV-A) limit its applicability. In EBR, throughput increases with group size but remains limited by the slowest group  $G_1$  (note that the y-axis in Figure 12b is a log plot). MassBFT achieves the highest throughput with stable latency.

**Nodes with different bandwidths.** MassBFT can tolerate the presence of slow nodes within a group. As shown in Figure 14, all nodes initially have a bandwidth of 40 Mbps. We gradually increase the number of 20 Mbps nodes in each group to study the corresponding changes in throughput and latency. When the number of 20 Mbps nodes exceeds 4, the throughput decreases by 36.9%, because 5 or more slow nodes cannot be treated as crashed. According to the transfer plan (Algorithm 1), in the best case (i.e., all slow senders send their chunks to slow receivers), log replication requires only 3 correct nodes out of 7 in each group. The latency also decreases by 13.4%, as entry replication becomes the bottleneck instead of execution.

### D. Scalability

Figure 13 presents the throughput results in the nationwide cluster under YCSB-A workload. To show the scalability of encoded bijective log replication, Figure 13a compares MassBFT with Baseline by scaling the number of nodes in each group from 4 to 40 (where  $f$  increases from 1 to 13). As the number of nodes per group increases, the throughput of Baseline decreases, as the group leader must send an entry  $f + 1$  times to each group. Since the group leader has limited WAN bandwidth, this increased redundancy reduces the overall throughput. In contrast, MassBFT utilizes the bandwidth of all nodes to replicate entries. As the number of nodes per group increases, the aggregate bandwidth of the group also increases. Consequently, the throughput increases accordingly. When the number of nodes per group exceeds 16, throughput remains stable because transaction signature verification during local PBFT consensus becomes the bottleneck.

We also evaluate the scaling performance by increasing the number of groups. We deploy 7 nodes in each of four additional data centers (Shenzhen, Beijing, Shanghai, and Guangzhou in China) and compare MassBFT with Baseline. In Figure 13b, when the number of groups scales from 3 to

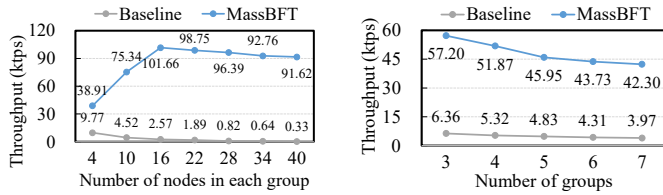


Fig. 13. Scaling performance (YCSB-A).

7, the throughput of MassBFT decreases from 57.20 ktps to 42.30 ktps, representing a decrease of **26.0%**. In contrast, the throughput of Baseline decreases from 6.36 ktps to 3.97 ktps, representing a decrease of **37.6%**. This is because the global Raft consensus is not scalable [57] [58], and increasing the number of groups leads to higher replication overhead.

### E. Fault Tolerance

To simulate node failures, each group contains two Byzantine nodes, and Byzantine nodes of all groups collude to replicate a tampered entry. To simulate data center crashes, we simultaneously kill all nodes in a group. Figure 15 shows the throughput and latency results in the nationwide cluster.

**Node Failures.** All Byzantine nodes always strictly follow the local consensus process and possess the same tampered entries. From the 20th second (the red dotted line), the Byzantine sending nodes encode these tampered entries (instead of the correct ones) into chunks and begins the encoded bijective log replication. Meanwhile, the Byzantine receiving nodes broadcast the chunks encoded from the tampered entries within its group. Additionally, all chunks sent by Byzantine nodes are received by the correct nodes. Figure 15 shows that the throughput of MassBFT remains unchanged because the correct nodes refuse to receive from the Byzantine nodes after rebuilding the tampered entries. As a result, a correct node can only receive chunks from other correct nodes to rebuild the entry, causing about a 3 ms increase in latency.

**Group Failures.** In Figure 15, group  $G_0$  crashes at the 40th second (the red solid line). After it crashes,  $G_0$  cannot assign timestamps from its clock  $clk_0$  to the newly replicated entries. Since deterministic ordering requires the timestamp  $e.vts[0]$  of entry  $e$  to be set before execution (Algorithm 2), entries can only continue replication but cannot be ordered or executed. As a result, latency increases and throughput reduces. After a timeout occurs [12] [13], a new leader is elected and starts assigning timestamps from  $G_0$ 's clock  $clk_0$  to entries, resulting in a reduction in latency and an increase in throughput. Finally, the throughput remains lower because the crashed group cannot propose entries.

## VII. RELATED WORK

**Asynchronous Consensus.** Asynchronous consensus protocols [59]–[68] are leaderless and are designed to operate in asynchronous networks [69]. HBBFT [59] and Dumbo [60], [61] employ ABA [70] and MVBA [71] [72], respectively, to identify and exclude the entries proposed by slow nodes. [63] proposes the Mempool protocol Narwhal for entry replication and the asynchronous consensus Tusk for entry ordering.

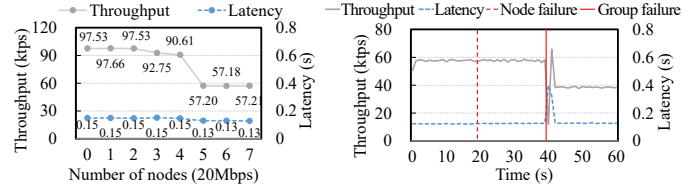


Fig. 14. Results varying node bandwidth (YCSB-A).

Fig. 15. Performance under failures (YCSB-A).

Narwhal stores entries in a DAG structure to maintain their causal order, enabling Tusk to establish a total order efficiently. Some protocols [66]–[68], adopt an optimistic fast lane to bypass ABA and MVBA. However, the fast lane is hard to achieve in practice [64]. Dumbo-NG [64] accelerates consensus by decoupling entry broadcast and ordering and leveraging pipelines, which is close to our work. However, it still suffers from high MVBA latency, which pipelining cannot hide. In contrast, MassBFT adopts Raft for global replication, which offers lower latency and tolerates  $\lfloor (n_g - 1)/2 \rfloor$  instead of  $\lfloor (n_g - 1)/3 \rfloor$  group failures since groups are crash-only.

**Erasure Coding.** In the CFT context, RS-Paxos [73] uses erasure coding [35] to reduce storage and network costs, but has limitations on the number of faulty nodes it can tolerate. [74]–[77] adopt different encoding and transmission strategies depending on the number of correct nodes available. In the BFT context, [78]–[82] use carefully designed encoding schemes to alleviate the leader bandwidth bottleneck. [40] proposes the delayed-replication algorithm to reduce the communication cost of notifying learners. RapidChain [83] and [84] integrate gossip with erasure codes to improve liveness in large networks. HBBFT [59] and Dumbo [60] integrate erasure codes into the reliable broadcast protocol [85], reducing network traffic during replication. BEAT [86] and ABFT [87] use optimized erasure codes to save bandwidth further. In addition, [88]–[94] apply erasure codes to reduce costs for BFT storage.

These protocols adopt a flattened architecture without fully utilizing the heterogeneous networks within and between data centers. In contrast, MassBFT adopts a hierarchical architecture that integrates erasure coding into group sending, eliminating leader bottlenecks while significantly reducing WAN traffic.

## VIII. CONCLUSION

This paper introduces MassBFT, a geo-distributed BFT consensus protocol designed for high performance and scalability. We propose encoded bijective log replication to reduce transmission redundancy, and asynchronous log ordering to enable groups to propose entries at their own pace without synchronizations. Experiments show that MassBFT outperforms SOTA protocols by a factor ranging from 5.49 to 29.96.

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