Moving Participants Turtle Consensus*

Stavros Nikolaou and Robbert van Renesse

Department of Computer Science, Cornell University, Ithaca, US
{snikolaou,rvr}@cs.cornell.edu

Abstract

We present Moving Participants Turtle Consensus (MPTC), an asynchronous consensus protocol for crash and Byzantine-tolerant distributed systems. MPTC uses various moving target defense strategies to tolerate certain Denial-of-Service (DoS) attacks issued by an adversary capable of compromising a bounded portion of the system. MPTC supports on the fly reconfiguration of the consensus strategy as well as of the processes executing this strategy when solving the problem of agreement. It uses existing cryptographic techniques to ensure that reconfiguration takes place in an unpredictable fashion thus eliminating the adversary’s advantage on predicting protocol and execution-specific information that can be used against the protocol.

We implement MPTC as well as a State Machine Replication protocol and evaluate our design under different attack scenarios. Our evaluation shows that MPTC approximates best case scenario performance even under a well-coordinated DoS attack.

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1 Introduction

Most distributed systems today are designed to tolerate failures. Existing fault-tolerance methods typically assume that failures are rare. They are tailored to provide good performance when no failures occur but might perform poorly under failure scenarios. However, as shown in works like [11], such designs allow malicious adversaries to craft workloads and Denial-of-Service (DoS) attacks that can substantially degrade the performance of certain state-of-the-art fault-tolerance protocols. As such DoS attacks become more common, it is becoming increasingly important to design fault-tolerance mechanisms that perform well in good scenarios while also gracefully handle adversarial ones. A core building block of many of these mechanisms are consensus protocols used by a set of replicas to agree on some state. One way to improve existing fault tolerance solutions is by enhancing the underlying consensus protocols with reconfiguration capabilities that allow them to change their execution parameters on the fly in order to better deal with adversarial workloads.

Our prior work on Turtle Consensus [22] also aims at attack-tolerant consensus. Turtle Consensus is a round-based consensus protocol that operates by using different consensus strategies across different rounds. The system’s processes try to reach agreement running a round of some consensus protocol in the literature; if they fail to do so they move onto the next round using a different protocol. The selection of each round’s strategy is predetermined and known to all processes running the protocol. The main strategy of Turtle Consensus is to use the best approach available for normal operation in a particular setting and switch to

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different “backup protocols” as soon as the approach becomes inefficient, for example in the case of a DoS attack. We showed that the approach used sub-optimal strategies during DoS attacks, thus bounding the protocol’s efficiency to the capabilities of these backup protocols. In addition, an adversary capable of compromising even a single consensus participant can learn and use the predetermined nature of the protocols’ succession to constantly drive the system to sub-optimal executions.

In this paper we address these concerns by adding another degree of freedom in the reconfiguration capabilities of Turtle Consensus. We present Moving Participants Turtle Consensus (MPTC), an extension to the Turtle Consensus protocol that allows switching not only the protocols but also the set of processes on which they run across different rounds of a single consensus instance. The consensus protocol round and the processes participating in its execution form what we call a configuration, which our approach changes unpredictably at each round. While the configuration selection for each round is predetermined by a trusted dealer, it is unknown to the processes during MPTC execution. Using existing cryptographic techniques, we ensure that, only if sufficiently many processes collaborate during some round, the next round’s configuration can be determined. This renders MPTC a valuable tool for building systems that can tolerate DoS attacks in both crash- and Byzantine-tolerant environments where a bounded portion of the system may be compromised.

2 Model

2.1 Processes and communication

Our system consists of a set of processes $\mathcal{N}$ that communicate using message passing. Each process is modeled as a state machine with a potentially unbounded set of states that executes transitions according to some protocol. The protocol specifies the transition function of the processes as well as the messages they exchange. Each process’s state consists of two components, the public and the private or secret state. The public state contains the description of the protocol that each process executes and any public cryptographic keys associated with the process. The secret state contains any run-time state the process manages during the execution of the protocol as well as any secret cryptographic keys and/or shares associated with the process.

Protocol execution and communication are asynchronous, meaning that there are no bounds on the time it takes processes to execute transitions and deliver messages.

A process can be correct or faulty. A correct process faithfully executes the protocol and is guaranteed to make progress as long as the conditions specified by the protocol at any given step are eventually met. In this paper we will primarily consider crash failures, although we extend our techniques to Byzantine failures in Appendix B. A faulty process may crash at any time after which it stops executing the protocol. Up to the point of the crash, processes faithfully follow the steps of the protocol. Communication between correct processes is reliable and secure. This means that, in the absence of a DoS attack (see below), messages sent by some correct process to another correct process are eventually delivered. It also means that messages are authenticated and cannot be tampered with or fabricated. We assume an upper bound, $f_c < |\mathcal{N}|$, on the total number of processes that might fail during the protocol’s execution.

Finally, we assume the existence of a special process $T \notin \mathcal{N}$ that from now on we will refer to as trusted dealer or simply dealer. The dealer is only used during initialization of the system during which it generates the initial public and secret state of all processes. We assume that during this setup phase the dealer is correct, that it can communicate via
secure channels with any process in the system, and that it does not disclose its state. After initialization, however, the dealer does not execute any protocol steps or exchange messages with any other process, and the dealer’s state is destroyed.

2.2 Adversary and attacks

We assume an adversary, $A$, that controls which processes fail and when. $A$ is limited on the number of processes it can fail by $f_c$ and cannot fail the dealer. $A$ can also control the delivery order of messages of all processes as well as delay communication, but must yield to the previously stated reliable communication assumption.

The adversary can also issue denial of service (DoS) attacks against the system that can fully saturate the bandwidth resources of at most $f_a < |N|$ correct processes. This can effectively prevent the targeted processes from progressing in the protocol’s execution since they can no longer communicate with the rest of the system. $A$ can change the targets of an attack over time and, in this way, can introduce communication and computation delays on certain processes. The adversary’s objective is to prevent the system from making progress. From now on we will denote by $f$ the maximum number of processes that can be crashed or under attack during the execution of the protocol, that is $f = f_c + f_a < |N|$.

In this work, we ignore DoS attacks that target other resources like CPU using legitimate traffic. These attacks can be mitigated using rate limiting techniques such as client cryptographic puzzles [2].

The adversary has read access to the public state of all processes as well as the secret state of up to $f$ processes. We call the processes whose secret state is disclosed to $A$ compromised.

While $A$ cannot modify this state, it can use this state to select the target processes of a DoS attack. Once $A$ has selected the set of compromised processes it can no longer change that set thus preventing $A$ from accessing the secret state of more than $f$ processes. Note that the set of compromised processes is not necessarily related to the set of processes that are crashed or under attack.

Finally, we assume that the various cryptographic schemes we are employing, like public key cryptography and threshold signatures, are secure in the random oracle model.

2.3 Cryptographic primitives

Our protocol relies on Threshold Coin Tossing [5]. Here we present a high-level description of this primitive that we will further formalize in Section 3. We employ an $(n, f + 1, f)$ threshold coin-tossing scheme in which $n$ parties maintain shares of an unpredictable function, $F$, mapping an arbitrary bit string, $r$, to a binary value $\{0, 1\}$. Each of these shares can be used along with an input $r$ to create a value that from now on we will refer to as function share.$^1$ At least $f + 1$ of these function shares of $r$ are required to reconstruct the result $F(r)$, while at most $f$ parties may be compromised. We will use the term function share of $F(r)$ to denote a function share of $r$ generated with a secret share of $F$.

The scheme defines three functions: 1) The split function, which takes as input a function $F$ (represented as a bit string) and creates a set of shares as well as a verification key for each of these shares. 2) The share combining function, combine, which takes an input $r$ of $F$ along with $f + 1$ valid function shares of $r$ and produces $F(r)$. 3) The share verification function verify, which takes an input $r$ of $F$, a function share on $r$, and the verification key.

$^1$ The term used in [5] for these values is coin shares
corresponding to the share that generated the input function share and determines whether
the function share is valid.

This scheme is based on threshold signatures [25] and can be used to create an unpre-
dictable sequence of bits while ensuring that it is computationally infeasible for the adversary
to produce an input $r$ and $f + 1$ valid function shares that once combined do not yield $F(r)$.
More formally, the scheme satisfies the following properties taken from [5]:

- **Robustness**: It is computationally infeasible for the adversary to produce a value $r$ and $f + 1$ valid shares of $r$ such that the result of the combine function is not $F(r)$.

- **Unpredictability**: Given a value $r$ and functions shares from fewer than $f + 1 - f$ correct processes, the adversary can predict the value of $F(r)$ with probability at most $\frac{1}{2} + \epsilon$ for negligible value $\epsilon \in \mathbb{R}$.

The previous unpredictability property was extended to sequences of output bits in [5],
such that, given a sequence of values $C_i$ for $i \in \{1, 2, \ldots, b\}$, an adversary with fewer than $f + 1 - f$ valid shares of some $C_i$ has negligible advantage in predicting $F(C_i)$. From now on, when we talk about unpredictability we will refer to this extended unpredictability property of threshold coin-tossing.

Note that the previously described extended unpredictability property allows us to share unpredictable functions in $\{0, 1\}^* \rightarrow \{0, 1\}^b$ for any finite $b$. In other words, we can model each such function as a random number generator that can produce $2^b$ different values and requires $f + 1$ processes to collaborate in order to produce the random (unpredictable) value corresponding to some arbitrary bit string $r$.

Threshold coin-tossing can be implemented using any non-interactive threshold signatures scheme that ensures unique valid signature per message as in [25]. A direct implementation of this scheme can be found in [5].

### 2.4 Underlying consensus protocols

MPTC, like other consensus protocols, solves the problem of agreement. In this problem,
a set of possibly distributed processes, each of which is initialized with some input value,
unanimously and irrevocably output one of those input values. More formally, let $\mathcal{N}$ be a set of processes each of which is initialized with some value from a value set $V$. Each process can employ either of the following primitives:

- **propose** a value which allows a process to communicate its value to the rest of the processes in $\mathcal{N}$,
- **decide** a value which allows a process to output a value.

Every correct consensus protocol must satisfy the following properties:

- **Validity**: If a process decides a value, then that value must be the input value of some process in $\mathcal{N}$.
- **Agreement**: If any two processes decide they must decide the same value.
- **Termination**: All correct processes eventually decide.

[13] has shown that in an asynchronous environment no consensus protocol exists that satisfies all of the above properties when even only a single failure can occur. To circumvent this result, a variety of protocols have been proposed [9] [10] that use a probabilistic approach and can guarantee the previous properties with the following modification on termination: All correct processes eventually decide with probability 1. For the remainder of this work we
will refer to the non-probabilistic description of termination as *definite termination* and to the probabilistic one as *probabilistic termination*.

A consensus protocol that implements the previous specification (using either definite or probabilistic termination) even under the presence of \(f\) crash failures is called \(f\)-crash-resilient. Note that our adversary can additionally perform denial-of-service attacks which can fully saturate a bounded number of processes and render them entirely unavailable. In an asynchronous environment there is no difference between a crashed process and a process that is under DoS attack from the other processes’ perspective. For this reason we say that a consensus protocol is correct in our model if it is \(f\)-crash-resilient where \(f = f_c + f_a\). From now on we will refer to such consensus protocols as \(f\)-resilient protocols.

Each process executing MPTC may run different consensus protocols at different rounds. We denote the set of possible protocols each process can choose from by \(\mathcal{P}\). Different consensus protocols make different assumptions under which they meet the previously described specification. The crash-tolerant consensus protocol of Ben-Or \[3\], for instance, assumes an asynchronous environment and that each infinite schedule has a bounded number of processes performing a finite number of steps. Other protocols make assumptions such as bounds on the number of failures, different degrees of synchrony, the existence of failure detectors \[8\], etc. We consider a consensus protocol correct if it satisfies agreement, validity and either definite or probabilistic termination. For each protocol \(P \in \mathcal{P}\), we denote the set of assumptions required to hold for \(P\) to be correct by \(A_P\). In other words, if assumptions \(A_P\) hold, then \(P\) satisfies validity, agreement, and termination. A protocol \(P\) is a valid candidate for \(\mathcal{P}\) if it is correct under both the assumptions in \(A_P\) and our previous model assumptions regarding failures, network reliability, and adversary.

We only consider consensus protocols operating in rounds and we follow the framework introduced in \[22\] for the specification of the round outcomes. According to this specification, every process running a round of a consensus protocol ends up in one of the following states: \(\{D, U, M\} \times \mathcal{V}\), where states \((D, v), v \in \mathcal{V}\) indicate that the process has decided \(v\), states \((U, v), v \in \mathcal{V}\) indicate that no process has decided up to the current round, and finally, states \((M, v), v \in \mathcal{V}\) indicate that while the process is not decided, if a decision was made by some process then it must have been \(v\). We will refer to these states as round outcomes or simply *outcomes*. We denote by \(o_r^p\) the outcome of process \(p \in \mathcal{N}\) at the end of round \(r \in \mathbb{N}\).

More formally the following invariants hold about the outcomes of processes completing a round of a correct consensus protocol in \(\mathcal{P}\):

1. **Invariant 1.** If \(\exists p \in \mathcal{N}, r \in \mathbb{N}\) such that \(o_r^p = (D, v), v \in \mathcal{V}\), then for each correct \(q \neq p \in \mathcal{N}\) it holds that \(o_r^q = (M, v)\) or \(o_r^q = (D, v)\).

2. **Invariant 2.** If \(\exists p \in \mathcal{N}, r \in \mathbb{N}\) such that \(o_r^p = (U, v)\) for some \(v \in \mathcal{V}\) then \(\forall q \in \mathcal{N}, u \in \mathcal{V}: o_r^q \neq (D, u)\).

This framework facilitates the description of MPTC in the next section and can be used to describe most consensus protocols in literature, including \[3, 8, 18\].

**Problem** Our goal is to design a round-based consensus protocol that is correct under the previous system and adversary assumptions and that runs a different existing consensus protocol on a different set of processes each round. The selection of protocols and processes for each round must not be predictable by the adversary without the collaboration of correct processes. For the purposes of this work, unpredictability is as described in Section 2.3.
In this section we describe our Moving Participants Turtle Consensus (MPTC) protocol. MPTC is an $f$-resilient consensus protocol operating in rounds such that in each round a different subset of processes may run a different consensus protocol. We start with some preliminary definitions and notation and then describe the protocol.

3.1 Participants and participant sets
MPTC is run by all processes in $\mathcal{N}$. In each round, only a subset of $\mathcal{N}$ is actively running a consensus protocol from a set of correct consensus protocols, $\mathcal{P}$. Let $\mathcal{P}_f$ correspond to the minimum number of processes required to run each protocol in $\mathcal{P}$. As an example, let $\mathcal{P}$ consist of the Ben-Or \cite{BenOr1983} and One-Third \cite{Dolev1985} consensus protocols. The first one requires $2f + 1$ processes to solve the agreement problem tolerating up to $f$ crash failures while the second one needs $3f + 1$. Thus $\mathcal{P}_f = 3f + 1$. We assume that $|\mathcal{N}| \gg f$ and thus $|\mathcal{N}| > \mathcal{P}_f$ for most $f$-resilient consensus protocols.

In the remainder of this paper, we say that a process runs or executes a protocol in $\mathcal{P}$ when it executes a round of that protocol. We will refer to a process executing a protocol in $\mathcal{P}$ at some round of MPTC as a participant or an active participant of that round. Let $PS = \{S \subseteq \mathcal{N} : |S| = \mathcal{P}_f\}$ be the set of all possible subsets of $\mathcal{N}$ where each subset has size $\mathcal{P}_f$. We call each such set a participant set. A process may be a member of multiple participant sets. In each round $r$ of MPTC, only a single participant set, $S_r$, is active, that is executing a consensus protocol in $\mathcal{P}$. We assume that participants in each participant set of some round $r$, $S_r \in PS$, are ordered and denote the $i$-th participant in $S_r$ as $S^i_r$. The active participant set for each round is determined by $T$ during initialization, which we describe later in this section.

3.2 Configurations
Before describing the initialization procedure and the core of MPTC, we need to define an important concept that encapsulates the information required for a set of processes to run a consensus protocol. We define a configuration of MPTC as a tuple $(P, S) \in \mathcal{P} \times PS$. $P \in \mathcal{P}$ describes the consensus protocol to run along with its initialization parameters. To better understand the information contained in the initialization parameters, consider a protocol like Lamport’s Paxos \cite{Lamport2001} and the core consensus protocol he called Synod. In Synod, processes play multiple roles, such as proposers and acceptors. In that sense, $P$ needs to encapsulate not only the protocol under execution, e.g. Synod, but also information related to its initialization such as mapping of proposers and acceptors to processes. The participant set $S \in PS$ corresponds to the set of processes that shall execute the consensus protocol specified by $P$. Let the set of all possible configurations $\mathcal{C} = \mathcal{P} \times PS$. Our approach implements a multi-party computation scheme for an unpredictable mapping between natural numbers (rounds) and configurations. We omit details regarding how to represent $P$ since this is an implementation issue and does not affect our protocol. We assume that $|\mathcal{C}|$ is bounded.

3.3 Initialization and trusted dealer
We are now ready to describe the initialization of our protocol, how we are using $T$ to create an unpredictable sequence of configurations, and how the active participants of a round can compute the corresponding configuration.
$T$ is a special process that generates the configuration that each process in $\mathcal{N}$ starts with in the first round. It also provides the processes the means to generate configurations for subsequent rounds. To achieve this, $T$ employs a $(\mathcal{P}_f, f + 1, f)$ threshold coin tossing scheme like the one described in Section 2.3. Using this scheme, $T$ shares a function $F_S$ between the $\mathcal{P}_f$ processes of each participant set $S \in \mathcal{P}$. Recall that threshold coin-tossing can be implemented using threshold signatures, thus when we say that $T$ shares a function $F_S$ with each participant set, in reality it simply selects a different public-secret key pair for each $S \in \mathcal{P}$ and shares the secret key. Given some round number $r$, at least $f + 1$ processes in $S$ need to collaborate to produce $F_S(r)$ while up to $f$ of them may get compromised. $f + 1$ is both a sufficient and necessary number of processes to compute the result of the function shared. $T$ cannot be compromised, failed or attacked by the adversary. At a high level, $T$ operates as follows:

1. For each $S \in \mathcal{P}$ the dealer picks a function $F_S : \{0, 1\}^* \rightarrow \mathcal{C}$ and generates a secret share, $h^S_p$, for each $q \in S$.
2. $T$ picks a configuration $C_0 \in \mathcal{C}$.
3. $T$ distributes $C_0$ and shares to processes over secure channels. $\forall S \in \mathcal{P}$ each process $p \in S$ receives $h^S_p$ and $C_0$.

Observe that each function shared by the dealer maps arbitrary strings to configurations. This differs from the functions we defined in Section 2.3 which map arbitrary bit strings to bit strings of some finite length $b$. Since $\mathcal{C}$ is finite, there exists $b = \lceil \log_2 |\mathcal{C}| \rceil$ such that we can trivially obtain an onto function $\{0, 1\}^b \rightarrow \mathcal{C}$. Thus, the functions we need to share can be trivially obtained by the ones supported by the threshold coin-tossing scheme.

Note that, by this high-level algorithm, a process in $\mathcal{N}$ will receive multiple shares, one for each participant set it belongs to. The dealer selects each $F_S$ such that the output is computationally indistinguishable from a randomly chosen function.

We now discuss how $T$ generates the secret shares. Given model parameters $\mathcal{P}$ and $f$, $T$ generates a different set of secret key shares for each subset, $S \in \mathcal{P}$. Each such set of secret key shares implicitly defines a function $F_S$ mapping bit strings to configurations. We call this operation split and it is similar to the threshold coin-tossing dealer initialization described in 5. split can be implemented using Shamir’s secret sharing 24 $(\mathcal{P}_f, f + 1)$. Note that the implementation in 5 is based on verifiable secret sharing because they are considering Byzantine failures. In our model, processes cannot lie and messages cannot be tampered with. As a result, no verification is needed within this context. We extend our approach to Byzantine failures in Appendix B where we introduce the required verification functionality.

Given a secret share, $h^S_p$, of some function $F_S : \mathbb{N} \rightarrow \mathcal{C}$ and some input, $r \in \mathbb{N}$, process $p \in S$ can create a function share of $F_S(r)$ using the share generation function, $GFS : S \times \mathbb{N} \rightarrow \mathcal{F}$ where $\mathcal{F}$ is the space of valid function shares that can be generated given a share $h \in S$ and a natural number. We define, $F^p_S(r) = GFS(h^p, r)$. A straightforward implementation of $GFS$ can be derived from the signature share generation for threshold signatures 25.

We define the combine functions as:

$combine : \mathcal{F}^{f + 1} \times \mathbb{N} \rightarrow \mathcal{C}$

$combine$ works by receiving function shares of some function $F_S$ and some input, $r$ and outputting $F_S(r)$ which corresponds to a configuration. More formally, let

$F^q_S(r) = \{F^q_S(r) \in \mathcal{F} | \forall q \in Q, Q \subseteq S \text{ and } |Q| = f + 1\}$
be any set of $f + 1$ function shares of $F_S(r)$, i.e. $F_S^Q(r) \in F^{f+1}$. Then we have that:

\[
\text{combine}(F_S^Q(r), r) = F_S(r)
\]

### 3.4 Protocol description

We can now describe the operation of each process executing MPTC. MPTC is an $f$-resilient round-based consensus protocol in which each round is executed under a different configuration. Let $C_r = (S_r, P_r) \in C$ be the configuration used for round $r$, where $S_r \in PS$ is the active participant set and $P_r \in P$ the consensus protocol specification for that round. Let $o_p^r$ be the outcome of a process $p \in S_r$ running $P_r$ at round $r$. Let a special value $\perp \notin V \cup C \cup \{(D, M, U) \times V\}$ represent the value of an uninitialized variable.

We assume that all processes have common knowledge of $\mathcal{N}$, $f$, $\mathcal{P}$, $C$ as well as of the functions $GFS$ combine. Each process $p \in \mathcal{N}$ runs MPTC with its identifier and some value $x_p \in V$ as input and at any point in time maintains the following state:

- its current round number, $r_p$, initialized to 0;
- its proposal $prop_{al}_p$, initialized to $x_p$;
- the outcome of a round, $o_p$ representing $p$’s decision state and initialized to $\perp$ at the beginning of each round; and
- the current configuration $c_p$ describing the currently known active participant set and the consensus protocol the active participants execute; it is initialized to $C_0$, which is provided by $T$ during the initialization phase.
- the secret shares, $h_p^S$, $\forall S \in PS$ such that $p \in S$ provided by $T$ during initialization.

We have organized MPTC description in phases. Messages exchanged between processes carry the number of the phase, the id of the sending process, and the current round along with the payload. Messages are of the form $\langle$phase number, process id, round, \ldots$$. Each round, $r$, of MPTC works in the following 3 phases:

- **Phase 1**: Each process $p \in S_r$ runs a round of the consensus protocol specified by $C_r$. Let $o_p$ be $p$’s outcome for round $r$. If $o_p = (D, v)$, then process $p$ updates $prop_{al}_p = v$, decides $v$ and never updates $o_p$ and $prop_{al}_p$ again in any future round. If $o_p = (M, v)$, then $p$ updates $prop_{al}_p = v$. Regardless of $o_p$’s value, $p$ goes to Phase 2.

- **Phase 2**:
  - **Step 1**: Each $p \in S_r$ computes function share $F_S^P(r) = GFS(h_p^S, r)$ and sends a Phase 2 message $\langle 2, p, r_p, o_p^r, F_S^P(r) \rangle$ to all processes in $S_r$. Then $p$ waits for Phase 2 messages from $P_f - f$ processes in $S_r$. Once $p$ receives enough messages from some $Q \subseteq S_r$, it proceeds to Step 2.
  - **Step 2**: If $o_p = (U, *)$ where $*$ can be any value in $V$, then $p$ updates its proposal to a value $v$, selected arbitrarily from the outcomes contained in the received Phase 2 messages. It also updates $o_p = (U, v)$.
  - **Step 3**: Let $F_S^Q(r)$ be the set of function shares received from processes in $Q$. $p$ computes the configuration of the next round, $r + 1$, as $C_{r+1} = combine(F_S^Q, (r), r)$ and moves on to Phase 3.

- **Phase 3**: Each $p \in S_r$ sends a Phase 3 message $\langle 3, p, r_p, o_p, C_{r+1} \rangle$ to each process in $S_{r+1}$. $p$ updates its state: $r_p = r + 1$, $c_p = C_{r+1}$ and if it is still undecided, it updates its outcome, $o_p = \perp$. Each process $q \in S_{r+1}$ that receives Phase 3 messages with the same configuration value, $C_{r+1}$, from $P_f - f$ processes, updates its proposal as follows. Let $R$ denote the set of outcomes received:
Case 1: If \( \exists o \in R \) such that \( o = (D,v) \), then \( q \) updates \( proposal_q = v \), decides \( v \), sets its outcome \( o_q = (D,v) \) and never updates \( o_q \) and \( proposal_q \) again in any future round.

Case 2: If \( \forall o \in R \) it holds \( o = (M,v) \) for some \( v \in V \) then \( proposal_q = v \).

Case 3: Otherwise, \( q \) selects an arbitrary outcome \((\ast,v)\) \(\in R\) where \(\ast\) can be any value in \(\{M,U\}\) and updates \( proposal_q = v \).

Then \( q \) sets \( r_q = r + 1 \), \( c_q = C_{r+1} \) and if it is still undecided, it sets \( o_q = \perp \). Finally, it starts the next round.

Figure 1 shows a visualization of the previous round description. MPTC runs for an unbounded number of rounds and eventually reaches a state in which a decision is made and all correct processes can eventually learn this decision. Messages from old rounds, either delayed in the network or sent by slow processes, are ignored while messages from future rounds are queued to be processed when the receiver reaches that round. The correctness of the previous protocol is presented in Appendix A. In addition, in Appendix B we present a byzantine-tolerant version of the previous protocol along with its own correctness discussion.

### 4 Implementation

In this section, we describe a simple implementation of a non-byzantine version of MPTC as well as a state machine replication protocol we built on top of it. To implement MPTC we need to decide on the following parameters: the choice of protocol set \( P \), the set of possible configurations \( C \), the configuration selection functions \( F_S, \forall S \in PS \), generated by the trusted dealer, and the implementations of \( split, GFS \) and \( combine \) functions.

Our set of protocols, \( P \), contains only a single consensus protocol, a parameterized version of single decree Paxos [18] in which each round comes with a predetermined leader known to all active participants. Paxos tolerates \( f \) crash failures using \( 2f + 1 \) processes and under failure-free execution conditions, it can reach a decision within a single round-trip of communication. We assume the weakest failure detector, \( \diamondsuit_W \), presented in [7] which we implement using timeouts with exponentially increasing timeout periods. This way we ensure that there will be enough rounds executed by sufficiently many processes, which is critical for ensuring termination in our Paxos variant.

The timeouts mentioned above may cause certain processes executing our Paxos variant to exit a round without knowledge of the round’s decision. Such processes need to retrieve this knowledge from the rest of the processes. To avoid incurring another round of communication in our Paxos variant, we piggyback this decision state retrieval onto Phase 2 of MPTC. Timed out processes can use the set of outcomes received to update their proposal.

Our set of configurations is \( C = \{(S,P) | S \in PS \text{ and } |S| = 2f + 1\} \) where \( P \in P \) is the described Paxos variant. Observe that in contrast to prior work on Turtle Consensus [22] we use the same protocol across configurations. In Turtle Consensus, different configurations used the same \( 2f + 1 \) set of processes. As a result, the adversary could try to track the current leader within that set of processes even if the leader changed across different configurations.
Therefore, a competent adversary could eventually locate and force Turtle Consensus rounds to fail, which can lead to poor performance. For that reason, Turtle Consensus kept switching between a leader-based (Paxos) and fully decentralized (Ben-Or) consensus protocols across configurations to prevent the adversary from exploiting the leader vulnerability. A side-effect of that approach, however, was that by falling back to a less efficient protocol (Ben-Or) it only achieved sub-par performance compared to the graceful execution using only Paxos rounds. With MPTC we do not need to employ such tactics since the adversary now needs to scan through $|\mathcal{N}| \gg f$ processes before it can identify the leader of our Paxos configuration.

In the implementation that we evaluate in Section 5 we did not implement the Threshold coin-tossing scheme. We emulated it instead by assuming that all participant sets use the same unpredictable function given to all processes via a configuration file. This file defines a sequence of configurations, one for each round, that processes move to in a round-robin fashion. We emulate the restrictions that the cryptographic framework imposes on the adversary by assuming that only the processes involved in rounds $r$ and $r + 1$ can learn $C_{r+1}$ and only after Phase 2 of round $r$ completes.

The interested reader can find an actual implementation of Threshold coin-tossing in [5]. In that work, they used cryptographically secure hash functions modeled as random oracles to implement unpredictable functions as well as for the GFS function. They also used Feldman’s verifiable secret sharing [12] for split function, though in our non-byzantine case Shamir’s secret sharing [24] can be used instead. Finally, for combine they use Lagrange interpolation with coefficients the computed function shares.

For more details, see Appendix C.1.

4.1 MPTC-based state machine replication

We used the previous implementation of MPTC to build a SMRP, similar to the one described in [22]. While the components of the implementation are similar, their interactions are different. There are three sets of processes, the clients, the replicas $\mathcal{R}$, and the participants $\mathcal{N}$. The clients issue requests to the participants who order these requests and forward them to replicas. Replicas execute the received requests in the order established by participants and send the results back to participants who then forward them back to clients. Participants can additionally send reconfiguration messages to each other in order to update the configuration of the MPTC execution.

In greater detail, clients send uniquely identifiable requests to sufficiently many participants in order to ensure that at least one correct participant receives each request. The participants receive requests from clients and are responsible for ordering these requests and send them for execution to the replicas. Only one participant set can be active at any point in time. Any participant outside that set receiving a client request relays that request to the currently known active participant set. Active participants receiving client requests spawn MPTC instances, one for each request that needs to be ordered. Each instance has its own identifier and decided requests are ordered according to the identifiers of the MPTC instances that decided them. Clients can only communicate with participants and thus they are unable to launch DoS attacks on the replicas. MPTC is lazily instantiated for each slot and MPTC messages carry instance identifiers so incoming protocol messages are properly processed by the correct instance. If an instance has not yet been created, messages for that instance are queued and processed when it is created. Finally, there are at least $f + 1$ replicas, each of which maintains a copy of state of the service implemented by the SMRP. All replicas are initialized in the same state and execute the clients’ requests in the order determined by id of the consensus instance created by the participants for each request.
5 Evaluation

In this section we present an evaluation of MPTC using the SMRP protocol presented in Section 4.1. In Section 5.1 we present the experiment setup and in Section 5.2 the performance results of MPTC under different attack scenarios.

5.1 Setup

We implemented MPTC and the SMRP described in Section 4.1 using C++. Our testbed consists of 10 nodes in Emulab [26], each with 8 cores running at 2.4 GHz, with 64GB of memory. For our experiments we used $f = 1$. Two nodes where designated as replicas, six as participants, one as clients, and one as the attacker. Nodes are connected by 1Gbps switched Ethernet as shown in Figure 2. Note that clients and attacker can only connect to participants, while participants connect to both replicas and clients. This choice was made to disable the attacker from directly attacking the replicas of SMRP and thus degrading performance without attacking the consensus mechanism. All communication between participants takes place through Switch 1 in our topology. Switch 2 is only used for participant to replica communication. We do not allow participants to communicate through Switch 2 since this would prevent the attacker from saturating the participants’ bandwidth with respect to the MPTC execution. This would give MPTC an unfair advantage and would not showcase the benefits of its reconfiguration capabilities. All communication is over TCP/IP except for the DoS attack traffic, which is entirely UDP/IP. One of the two client nodes is used by the attacker and the other for creating legitimate client threads. We use a separate node for attacks in order to limit the effect of bandwidth attacks on the clients’ ability to issue requests.

To simplify our evaluation, we set $\mathcal{C}$ to contain only two configurations such that the corresponding participant sets are disjoint. The configuration selection function provided by the trusted dealer (in our implementation by a configuration file) simply alternates between these two configurations every time a round fails. The predetermined Paxos leader of each configuration depends on the round in which the configuration is run and is rotated in a round-robin fashion every time the same participant set is reused. We consider that the attacker does not have this knowledge to make informed decisions regarding targeting processes.

Clients first connect to $f + 1$ random participants to which they issues requests. Once connected, each client executes the following loop: It issues each request to all $f + 1$ participants, waits for a response, discarding duplicate responses, and then sends the next request. Note that by connecting to $f + 1$ participants, we ensure that each client request reaches at least one correct participant who will further forward the request to the active participants. We have client requests contain no-ops, which means that when a decided request becomes ready for execution, replicas can immediately reply with a response.

![Figure 2: Experiment topology.](image-url)
The attacker creates a small number of attack threads, each of which targets a single participant, selects a random port, and sends UDP dummy messages as fast as it can. Note that these messages are not requests and are not processed by our participants since they never get to the application level. As in the Turtle Consensus evaluation [22], the goal of the attack is to prevent at most one participant from participating in MPTC instances. The attacker can focus all threads on the same participant or spread them across different ones. Since all attack threads are created on a single node, the aggregate bandwidth the attacker threads can saturate from the service cannot exceed 1Gbps.

We conducted experiments to test the throughput and latency of our implementation under normal execution and DoS attacks. Both metrics were measured at the client side. For throughput we measured the aggregate number of operations per second completed by client threads. Note that this is not the actual number of instances completed per second by our SMRP implementation since the same request might be decided more than once.

Other parameters of our experiment include:

- Duration: Each experiment lasted 1 minute. We found longer experiments did not significantly affect our metrics.
- Load: The number of concurrent clients, which ranged in our experiments from 1 to 64.
- Request size: The size of the command contained in each client request, which we set to 100 bytes.
- Attack message size: The size of the UDP messages send by attack threads to saturate the participants bandwidth; we set that to 1KB since our experimentation with our platform showed it is the smallest message size with the best results for the attacker.
- Number of attacker threads: Each run involving a DoS attack had 8 attack threads. We found that this number of threads yields best results for the attacker even when all 64 clients are connected to the target sharing the same link.
- Timeout: This is the initial timeout period used in our Paxos variant (Section 4) for each MPTC instance. Every time a round of some instance fails we double the timeout period for that instance.

5.2 Results

In our evaluation, we investigated three main scenarios. In the first, we run our implementation of MPTC without any attacks taking place. The performance of this scenario will be our baseline since any attack scenarios drain resources from the system and thus is expected to perform similar or worse. This scenario is labeled “No attacks” in our figures.

The second scenario has the attacker focusing the DoS attack on a single node, the one that hosts the Paxos leader; we set that to 1KB since our experimentation with our platform showed it is the smallest message size with the best results for the attacker.

In our figures, this scenario is labeled “Attack leader without reconfiguration”.

Finally, the third scenario uses an attacker who like in the previous scenario focuses on a single node. In this scenario the attacker is given the initial position of the leader but this time our implementation uses the MPTC version we described in Section 5.1, where consensus instances execution alternates between two disjoint sets of nodes. The attacker strategy...
here is to saturate the bandwidth of the known leader. It keeps attacking that node for the entirety of the experiment run. This attack is labeled “Attack leader with reconfiguration”.

Figure 3 Moving Participants Turtle Consensus performance under different attack scenarios. Figure 3a shows the throughput comparison of the previous three experiment scenarios as a function of the load on the SMRP. Each point represents the average throughput over 10 runs for each number of clients. In each of these runs clients connect to random participants, which in turn means that performance will vary across experiments. The first scenario is our best case scenario since the system operates at full resource capacity. The second scenario shows that performance suffers substantially when the Paxos leader is under attack. This is to be expected since the leader’s participation is critical for making progress in each MPTC instance. In the third scenario we observe the benefits of the reconfigurable version of MPTC in action. The SMRP throughput is close to that of the No Attacks case. The main reason for this behavior is that since the leader of the first configuration is under attack and lacks the bandwidth to handle the valid traffic, some instance will inevitably fail the first round since the remaining participants will eventually time out. That will cause a reconfiguration that changes the active participant set. The new participants will pick up the failed instances as well as future requests and continue operating at full capacity. The minor deviations observed between scenarios 1 and 3 are mainly due to the randomness of client distribution over the set of all participants.

Figure 3b shows a comparison of the same scenarios as the load increases, but this time with respect to latency. Observe that all scenarios behave similarly with the latency linearly increasing with the load. This behavior is to be expected since, as the load increases, the number of concurrent MPTC instances increases, which in turn increases the latency for each client. After all, each of them has to wait for a response to their previous request before sending the next one. As in the case of throughput, we see that both scenarios 1 and 3 have similar latencies while scenario 2 performs poorly. The reasoning is the same. In the second scenario the leader under attack is slower in completing instances, which raises the wait time for each client.

Note that this evaluation does not take into account the additional cost of reconfiguration that stems from the cryptographic operations required for threshold coin tossing like RSA exponentiations. We therefore expect that under frequent reconfigurations there will be a wider gap between the performances of scenarios 1 and 3. However, we also expect that such reconfigurations will be infrequent, especially as the number of processes increases. Thus, while not an absolute comparison, our evaluation showcases the expected behavior and advantage of MPTC.
6 Related Work

A wide range of crash-tolerant consensus protocols have been proposed in literature each optimized for a different setting and/or metric. Some were designed to handle datacenter-scale systems like [6], which describes how Paxos was used to implement a fault-tolerant database for the Chubby locking service, an instance of which lies in each Google's datacenter. Others are focused on wide area deployments such as Mencius [21], which is a Paxos variant that employs multiple leaders each of which is responsible for a different set of consensus instances and may reside at different datacenters. Another important differentiating aspect of consensus protocols is whether they employ a special leader process like in [8, 18] or whether they are fully decentralized like the protocol proposed in [3]. This can greatly affect the behavior of a consensus protocol under different failure scenarios, including attacks, and was thus used by previous work on reconfigurable consensus [22] to design consensus protocols that provide acceptable performance under certain DoS attacks.

Our work resembles the work on Vertical Paxos [19]. Vertical Paxos is a reconfigurable state machine replication protocol that uses a special auxiliary master process to decide the next configuration of the system including the set of replicas participating in that configuration. Unlike Vertical Paxos, MPTC does not require additional master processes to be constantly active in order to compute the next configuration. Our assumed trusted dealer is only active during initialization. In addition, Vertical Paxos is not designed for an adversary capable of compromising even a single process and thus would not perform as well against the DoS attacks described in this work.

Moving target defenses have often been used as response to DoS and Distributed DoS (DDoS) attacks. [14] proposes changing the IP address of the target node for dealing with local IP-based DoS attacks. More recently in [16], Software-Defined Networking (SDN) has been used to implement moving target defense approaches like “random host mutation” in which, similarly to [14], the controller periodically alters the virtual IP addresses of hosts to hide the real IP addresses from an intruder. Our Moving Participants Turtle Consensus approach resembles more the “proactive server roaming” approach in [17]. That is an adaptive approach in which the active server proactively switches servers from an existing pool in order to deal with unpredictable and undetectable attacks. Their approach ensures that only legitimate clients can track the moving server. Like in the case of our MPTC protocol, proactive server roaming performs gracefully during attacks. However, it imposes significant overhead in attack-free scenarios, which is not the case for MPTC since we only reactively change configurations.

Our work assumes an adversary that cannot change the set of corrupted processes over time. Other related work has focused on dynamic models of corruption. [15] introduced proactive secret sharing, an instance of proactive security [23] for supporting secure computation in synchronous distributed systems. These ideas have been adapted to asynchronous ones in [4, 27]. While these approaches did not consider DoS attacks, they are orthogonal to ours and can be used to further improve this work for dealing with mobile adversaries.

Running consensus on a subset of a larger set of processes to decrease message complexity has been explored in [1]. It has also been explored more recently in [20] for improving the scalability of Byzantine agreement on blockchains.

7 Conclusions

In this paper we presented Moving Participants Turtle Consensus (MPTC), an extension to the Turtle Consensus protocol [22] that allows running different consensus protocols, on
different sets of processes, across different rounds of a single consensus instance. MPTC can deal with adversaries with bounded information on the system by making unpredictable changes in the execution of the protocol. Our evaluation of our prototype implementation of MPTC suggests that we can achieve the performance offered by the most efficient consensus protocols even when the system is under attack.

There are various directions for further exploration and improvement of MPTC. First, MPTC should be tested under more sophisticated attacks, for example in which the attacker keeps changing the target of the attack even after reconfigurations, until it hits the node whose absence most impacts performance. In addition, we would like to design a dynamic configuration selection scheme in order to extend MPTC’s applicability to different environments where unexpected changes in the workload of the system may lead to sub-optimal configurations.

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MPTC needs to satisfy the correctness properties of consensus protocols we defined in Section 2. We will show these properties assuming that all the underlying protocols used in \( \mathcal{P} \) are \( f \)-resilient.

Recall that each protocol \( P \in \mathcal{P} \) potentially makes additional assumptions regarding the system model under which the protocol satisfies the previous properties. To guarantee correct execution of \( P \) under MPTC, the protocol must be correct under the previous set of assumptions and those assumptions made in Section 2. Note that this union may contain different assumptions about the same property or aspect of the system. Let \( P \) be a consensus protocol for synchronous systems. Then this union of assumptions will contain our assumption that the system is asynchronous and \( P \)'s assumption that it is synchronous. In such cases, where one assumption is stronger than another with respect to a particular aspect of the model, we assume that the stronger assumption holds. In our example that would mean that the system would be synchronous.

To facilitate our discussion we will introduce some notation to describe collections of assumptions for the protocols in \( \mathcal{P} \). Let \( \mathcal{A} \) denote the set of assumptions that we made in Section 2. For each \( P \in \mathcal{P} \), let \( \mathcal{A}_P \) denote the set of assumptions \( P \) makes about the system and its operation in order to tolerate \( f \) crash failures while satisfying the correctness criteria stated in Section 2. Such assumptions may regard the minimum number of processes required as a function of \( f \), the behavior of the network, the time and ordering of events, restrictions on the adversary, and many more. In this work we will not attempt to accurately model such assumptions since their descriptions vary greatly among different consensus protocols proposed in literature.

Given \( \mathcal{A} \) and \( \mathcal{P} \) we define the set of assumptions that must hold in every execution of MPTC as

\[
\mathcal{A}_P = \left( \bigcup_{P \in \mathcal{P}} \mathcal{A}_P \right) \cup \mathcal{A}
\]

In other words, every execution of MPTC must respect the union of all assumptions made by every individual protocol used in Phase 1 as well as our model’s assumptions. If an assumption is overridden by another, the stronger of the two holds and thus weaker assumptions are excluded from \( \mathcal{A}_P \). Under this definition of \( \mathcal{A}_P \) we have that \( \forall P \in \mathcal{P} \) every execution of \( P \) under \( \mathcal{A}_P \), respects the correctness properties of consensus protocols.

We will first discuss validity and agreement and finally we will argue about termination.

### A.1 Validity

In order to satisfy validity, MPTC needs to ensure that any value decided must be the input value \( x_p \) of some process \( p \in \mathcal{N} \). This is encapsulated in the following lemma:

\textbf{Lemma 1.} If a process in \( \mathcal{N} \) running MPTC under \( \mathcal{A}_P \), decides a value \( v \in \mathcal{V} \), then \( \exists p \in \mathcal{N} : x_p = v \).

\textbf{Proof.} (Sketch) A process running MPTC can decide during Phase 1 or Phase 3. Any decision made during Phase 1 respects validity by correctness of each protocol in \( \mathcal{P} \) under \( \mathcal{A}_P \). By the outcome and proposal update function of Phases 2 and 3 we described in Section 3.4 the processes’ proposals and thus possible decision values can only come from the values of the outcomes produced in Phase 1. Since validity holds during Phase 1, if some value \( v \) gets decided at some round \( r \), \( \exists p \in \mathcal{N} \) whose input value in \( r \) is \( v \).
From the previous lemma, we have that MPTC initialized with a valid \( P \) satisfies validity under \( A_P \).

### A.2 Agreement

Agreement is a safety property that ensures that no bad states occur during the system’s operation. All correct consensus protocols must satisfy the agreement property for the entirety of their execution. Thus we need to show:

**Lemma 2.** If any two processes running MPTC under \( A_P \) decide values \( v \) and \( v' \) respectively, then \( v = v' \).

**Proof.** (Sketch) Assume that \( \exists p, p' \in N \) that decide values \( v, v' \in V \) in rounds \( r, r' \in \mathbb{N} \) respectively such that \( v \neq v' \). We have the following two cases:

- **Case** \( r = r' \): For two different processes to decide different values it must be the case that outcomes \( (D, v) \), \( (D, v') \) are computed after Phase 1 of \( r \). This cannot occur because each \( P \in \mathcal{P} \) is correct under \( A_P \) and therefore respects agreement.

- **Case** \( r \neq r' \): W.l.o.g. assume \( r < r' \). There must be a round \( \bar{r} \) such that \( r \leq \bar{r} < r' \) in which all correct processes in \( S_{\bar{r}} \) compute either \( (D, v) \) or \( (M, v') \) outcomes during Phase 1 and after which \( \exists q \in S_{\bar{r}+1} \) such that either \( a_q = (D, v') \) or \( a_q = (M, v') \) and \( v' \neq v \). This is impossible: First, Phases 2 and 3 of round \( \bar{r} \) only allow processes to change their values to those of the outcomes computed during Phase 1 of \( \bar{r} \), which in our scenario will be \( v \). In other words, all values received by processes in \( S_{\bar{r}+1} \) in Phase 3 of round \( \bar{r} \) will be \( v \). Second, by the validity property of the protocols in \( \mathcal{P} \) no outcome \( (D, v') \) or \( (M, v') \) can be computed after Phase 1 of round \( \bar{r} + 1 \) if all processes start with proposal \( v \).

By contradiction it must hold \( v = v' \). ▷

### A.3 Termination

Termination encapsulates the liveness or progress requirement on consensus protocols by ensuring that eventually a decision is or can be made. In our model (Section 2), we stated two versions of termination, the definite and the probabilistic one. Notice that by definition, definite termination implies probabilistic termination.

Recall that in MPTC each consensus round is run by a subset of the system’s processes. It is possible that in an infinite execution some correct processes may only execute a finite number of consensus protocols rounds and as a result not be able to decide. Note that even if at least one correct process decides, all correct processes can eventually learn this decision by having the decided processes broadcast a special decision message to all processes in \( N \), which in turn decide upon reception of that message. Therefore to guarantee definite (probabilistic) termination we simply need to show that eventually at least one correct process in \( N \) decides (with probability 1).

MPTC termination depends on both the guarantees provided by protocols in \( \mathcal{P} \) and the properties of the interleavings of rounds of different protocols during MPTC execution. The guarantees of each \( P \in \mathcal{P} \) allow us to reason about the properties that each MPTC round satisfies. We know that every round of \( P \) must produce an outcome in \( \{D, M, U\} \times V \) provided that \( A_P \) hold. If not, \( P \) would violate termination. Note that \( P \) should also satisfy termination under \( A_P \) as well, since \( A_P \) makes the same or stronger assumptions. If \( P \) guarantees definite termination, then in any infinite execution of \( P \) there must be at least one correct process \( p \in N \) that produces outcome \( (D, v), v \in V \) in infinitely many rounds. If \( P \) guarantees probabilistic termination then \( \exists \epsilon \in (0, 1] \) such that in any infinite execution
of \( P \), infinitely many rounds have probability at least \( \epsilon \) for at least one correct process to produce outcome \( (D, v), v \in V \).

Given the previous guarantees provided by the protocols in \( P \) we can show the following:

▶ **Lemma 3.** Every correct process executing a round of MPTC under \( A_P \) eventually completes that round.

**Proof.** To show the above we need to ensure that each active participant of some round \( r \) completes the all three Phases. Phase 1 completes by the termination property of each protocol in \( P \) with each correct process in \( S_r \) computing an outcome with respect to the protocol specified by configuration \( C_r \). Phases 2 and 3 rely on each correct process eventually receiving \( P_f - f \) messages which will occur due to our assumptions on network reliability and maximum number of failures and processes under attack.

Lemma 3 ensures no process participating in any of the MPTC Phases ever blocks. This *non-blocking* property of MPTC rounds, however, is not sufficient to ensure progress. To reason about progress, we need to reason about the effect of interleaving rounds of different protocols on infinite executions.

To reason about interleavings of consensus protocol rounds we need to reason about configuration sequences and thus about the properties of the unpredictable functions \( F_S \) selected by dealer \( T \) for each participant set \( S \) during initialization. Consider the case where \( P \) contains two artificial protocols \( P, P' \) such that any process running \( P \) can only decide during an odd round, while any process running \( P' \) can only decide on an even round. Assume a pathological infinite execution in which \( C_r = (P, S_r) \) if \( r \) is even and \( C_r = (P', S_r) \) if \( r \) is odd. If we can define \( F_S \) for each round \( r \) in such a way that the previous configuration sequence is generated during Phase 2 of MPTC, termination cannot be achieved.

More formally, let the set of possible configurations, \( C \), be based on a valid \( P \). We define a finite sequence \( I \) of configurations in \( C \) as *conforming* if executing MPTC under \( A_P \) such that \( I \) appears infinitely often, at least one correct process computes a decision outcome in infinitely many rounds. Any finite sequence of configurations in \( C \) that does not have the previous property is called *non-conforming*. The case described in the previous paragraph is an example of such a non-conforming sequence. From now on we will refer to finite sequences of configurations in \( C \) as interleavings.

The previous notion of conforming interleavings raises another restriction on any implementation of MPTC and more specifically on the choice of \( P \). In every implementation there must be at least one conforming interleaving containing configurations of \( C \). We call an infinite execution of MPTC in which the corresponding configuration sequence contains infinitely many conforming interleavings, a *conforming execution*.

In the random oracle model, which we assume in this work, the configurations sequences generated by the shared functions \( F_S \) have the following property:

▶ **Lemma 4.** Let \( C \) be a set of configurations based on a valid set of protocols \( P \) and let \( F_S \) be an unpredictable function for each \( S \in PS \) generated by \( T \). Assuming a conforming interleaving \( I \) exists in \( C \), any infinite sequence of configurations corresponding to an infinite execution of MPTC contains infinitely many occurrences of \( I \).

**Proof.** Each \( F_S \) used to generate the next configuration in Phase 2 of each MPTC round is based on the threshold coin-tossing mechanism described in Section 2. The unpredictability of this mechanism is based on the use of cryptographic hash functions. In the random oracle model, given some input \( r \) these functions produce a value chosen uniformly at random from their co-domain. In other words, at each round \( r \) there is a positive probability for each
configuration \( C \in \mathcal{C} \) to be selected as the next configuration to run. Therefore in an infinite execution, any interleaving in \( \mathcal{C} \) appears infinitely often. Thus conforming interleaving \( \mathcal{I} \) appears infinitely often.

By the previous Lemma we have that:

▶ **Corollary 5.** Assuming a conforming interleaving \( \mathcal{I} \) exists in \( \mathcal{C} \), any infinite execution of MPTC is conforming.

We can now reason about the termination guarantees of MPTC, under the assumptions that \( \mathcal{P} \) is valid, that there is a conforming interleaving in \( \mathcal{C} \) and that \( \mathcal{AP} \) hold during MPTC’s execution. Depending on the protocols executed under the configurations of a conforming interleaving, definite or probabilistic termination can be guaranteed. This is shown in the following lemmas:

▶ **Lemma 6.** Let \( \mathcal{I} \) be a conforming interleaving in \( \mathcal{C} \) that contains at least one round of a protocol satisfying definite termination. In any infinite execution of MPTC there is at least one correct process that makes a decision.

**Proof.** By Lemma 4 we have that \( \mathcal{I} \) will be executed infinitely often and so will the round of some protocol \( P \) satisfying definite termination. By the termination guarantees of \( P \) it must be the case that at least one correct process computes a decision outcome \((D, v)\) for some \( v \in \mathcal{V} \) in infinitely many rounds. Therefore eventually at least one correct process running MPTC decides.

We can state a similar lemma for probabilistic termination:

▶ **Lemma 7.** Let \( \mathcal{I} \) be a conforming interleaving in \( \mathcal{C} \) that contains rounds of protocols satisfying probabilistic termination. In any infinite execution of MPTC there is at least one correct process that makes a decision with probability 1.

**Proof.** The proof is similar to that of Lemma 6 except from the fact that the rounds in which a decision can be made can only ensure that at least one process decides with probability 1.

Finally, we need to show the security properties satisfied by MPTC under the previous assumptions. More specifically we need to show the following two properties:

▶ **Lemma 8.** Robustness: It is computationally infeasible for \( \mathcal{A} \) to produce \( r \) and \( f + 1 \) shares of \( r \) for any participant set \( S \) such that the output of combine given the previous shares and value as input is \( F_S(r) \).

**Proof.** It follows directly from the robustness property of the threshold coin-tossing scheme we use for computing the next round’s configuration in each MPTC round. The property is proven in [5].

▶ **Lemma 9.** Unpredictability: Let \( C^A_r \) be \( A \)’s prediction for \( F_{S_r}(r) \) after learning at most \( f \) function shares for \( F_{S_r}(r) \) as well as any number of functions shares for \( F_{S_{r'}}(r') \) for arbitrary many \( r' < r \). Then the probability of \( C^A_r = F_{S_r}(r) \) is at most \( \frac{1}{2^{m}} + \epsilon \) where \( \epsilon \in \mathbb{R} \) is negligible.

**Proof.** This property follows from the implementation of each \( F_S \) as a \((P(f), f + 1, f)\) threshold coin-tossing scheme and from the extended unpredictability property of a sequence of coins produced by this scheme shown in [5]. By selecting the length of the sequence to be \( m = \lceil \log |\mathcal{C}| \rceil \) we ensure that the probability of \( A \) predicting the next configuration having compromised at most \( f \) processes in the active participant set is \( \frac{1}{2^{m}} + \epsilon \leq \frac{1}{|\mathcal{C}|} + \epsilon \) for negligible security parameter \( \epsilon \).
The previous termination discussion relies on the random oracle assumption we made earlier. While this assumption is important for supporting the unpredictability and robustness properties of our configuration generation scheme it is not necessary if such properties are not needed. If we wanted to drop this assumption, we would need to place additional restrictions on \( P \) and \( C \) to satisfy termination for MPTC. More specifically, we need to ensure that any interleaving in \( C \) is conforming. Under that assumption the termination results still hold since any infinite execution consists of conforming interleavings in which at least some correct process makes a decision.

\section*{B Extension to Byzantine Failures}

\subsection*{B.1 Byzantine agreement model}

Our MPTC protocol can be extended to support byzantine agreement. Under this weaker adversary assumption, we call a process honest if it faithfully executes the protocol. An honest process may crash during the execution but up to the point of crash its execution does not deviate from the protocol description. Observe that in our previous model (Section 2) all processes where honest. We call a process correct if it is honest and eventually makes progress. This implies that a correct process never crashes. Faulty processes, on the other hand, may deviate arbitrarily from the protocol but cannot alter the secret state.

The specification of the protocols in \( P \) also change. Each protocol in \( P \) is now a Byzantine Fault-Tolerant (BFT) agreement protocol. In this problem, the objective is for all honest processes to agree on the same value. The correctness criteria are as follows:

- **Agreement**: If two honest processes decide, they decide the same value.
- **Validity**: If an honest process decides value \( v \), then \( v \) was proposed by at least some process.
- **Unanimity**: If all honest processes propose the same value \( v \), then an honest process that decides must decide \( v \).
- **Termination (Definite)**: All correct processes must eventually decide.
- **Termination (Probabilistic)**: All correct processes must eventually decide with probability 1.

Note that it now holds \( \mathcal{P}_f \geq 3f + 1 \). Also note that the byzantine failures assumption is now part of our model assumptions set \( \mathcal{A} \) and thus is included in \( \mathcal{A}_P \) for any set of BFT protocols \( \mathcal{P} \). The round outcomes framework described in Section 2 still hold under the following modifications on invariants 3 and 4:

\begin{itemize}
  \item **Invariant 3.** If there exists honest \( p \in \mathcal{N}, r \in \mathbb{N} \) such that \( o_p^r = (D, v) \) where \( v \in \mathcal{V} \) then for each correct \( q \neq p \in \mathcal{N} \) it holds that \( o_q^r = (M, v) \) or \( o_q^r = (D, v) \).
  \item **Invariant 4.** If there exists honest \( p \in \mathcal{N}, r \in \mathbb{N} \) such that \( o_p^r = (U, v) \) for some \( v \in \mathcal{V} \) then there is no honest \( q \in \mathcal{N}, u \in \mathcal{V} \) such that \( o_q^r = (D, u) \) and \( u \neq v \).
\end{itemize}

This new version of invariants refer to honest processes since faulty processes may update their state arbitrarily at any point in time. Thus the processes’ outcomes are meaningful only for honest and correct processes.

\subsection*{B.2 Trusted dealer protocol}

Let \( \mathcal{P} \) be a set of BFT protocols that tolerate up to \( f \) failures. The trusted dealer initialization protocol now becomes:

\[ \text{...} \]
1. The dealer assigns an identity for each \( p \in \mathcal{N} \) using a public-key cryptography scheme. It generates a public-private key pair, \((\text{public}_p, \text{private}_p)\) for each \( p \in \mathcal{N} \).

2. For each \( S \in \mathcal{P}_S \) the dealer picks a function \( F_S : \mathbb{N} \to \mathcal{C} \) and generates a verification key \( VK_S \) and for each \( q \in S \) a secret share, \( h^q_S \) and a verification key \( VK^q_S \).

3. \( T \) distributes shares and keys to processes over secure channels. \( \forall S \in \mathcal{P}_S \) each process \( p \in S \) receives \( h^p_S, VK_S, \) and \( VK^q_S, \forall q \in S \). In addition, each process \( p \) receives \((\text{public}_p, \text{private}_p)\) and the public keys of all other processes.

In the byzantine case, we need to use a verifiable secret sharing scheme like [12] since we need a way to ensure that invalid function shares created by faulty processes can be identified and discarded by honest ones. In such schemes, a verification function is specified which typically works by receiving a value \( r \), a share of \( F_S(r) \) and some verification keys that depend on \( F_S \) and the secret share used to generate the previous function share and outputs 1 or 0 indicating whether the share provided is a valid share of \( F_S(r) \) or not. We define our verification function, \( \text{verify} \), as follows:

\[
\text{verify} : \mathbb{N} \times \mathcal{F} \times \mathcal{K}^2 \to \{0, 1\}
\]

where \( \mathcal{K} \) is the space of verification keys. Note that the \( \text{verify} \) above is modeled after the share verification algorithm presented in [5]. Given a function share \( F_S^q(r) \), it receives two verification keys, \( VK_S \), and \( VK^q_S \in \mathcal{K} \), the first produced using \( F_S \) and the second using \( h^q_S \).

\( T \) can generate these verification keys using a secure cryptographic hash function, that is a function that is easy to compute but computationally infeasible to reverse. Feldman [12] provided some example functions with this property, with RSA being one of them. We abstract away such details and denote by \( \text{hash} : \{0, 1\}^* \to \mathcal{K} \) a function that has this property. \( T \) can use \( \text{hash} \) to generate the above mentioned verification keys during phase 1 (after using split) as follows:

\[
VK_S = \text{hash}(F_S), \forall S \in \mathcal{P}_S \\
VK^q_S = \text{hash}(h^q_S) \forall p \in S, \forall S \in \mathcal{P}_S
\]

Note that our \( \text{hash} \) function takes different types of input, such as \([\{0, 1\}^* \to \mathcal{C}]\) and \( \mathcal{S} \), but both are bit strings and so are the elements of \( \mathcal{K} \) and \( \mathcal{F} \). While we are using different domain notations to distinguish between functions, secret and function shares, and verification keys, any implementation of these schemes is working with bit strings.

### B.3 Byzantine Tolerant MPTC

The Byzantine version of our MPTC protocol is similar to that in Section 3.4. The processes maintain the same state as in Section 3.4 plus the cryptographic keys generated by the dealer. All messages are signed using these keys. Messages are of the form \((\text{Phase number, process id, round, signature, ...})\) and processes ignore messages with invalid signatures or messages not destined to them. The protocol operates in rounds, the configuration of the first round, \( C_0 \), is determined by \( T \) and is known by all processes and each round proceeds as follows:

- **Phase 1:** Each process \( p \in S_r \) runs a round of the BFT protocol specified by \( C_r \). Let \( o_p \) be \( p \)'s outcome for round \( r \). If \( o_p = (D, v) \), then process \( p \) updates \( \text{proposa}l_p = v \), decides \( v \) and never updates \( o_p \) and \( \text{proposa}l_p \) again in any future round. If \( o_p = (M, v) \), then \( p \) updates \( \text{proposa}l_p = v \). Regardless of \( o_p \)'s value, \( p \) goes to Phase 2.
The byzantine version of MPTC must satisfy the properties described above as well as have been proposed by some process. In fact, the termination arguments for the byzantine MPTC are exactly the same. Robustness and unpredictability are also exactly the same and our arguments are mainly borrowed from the corresponding arguments of the Byzantine agreement protocol of [5]. For the rest of the properties we have the following results:

Lemma 10. Byzantine MPTC satisfies validity.

Proof. By correctness of the underlying BFT, any decision made by an honest process during Phase 1 respects validity. A decision during Phase 3 can only occur if at least \( f + 1 \) processes send the same decision outcome \((D, v)\) for some \( v \in \mathcal{V} \). For that to happen at least one honest process must have computed that outcome during Phase 1. Thus that value must have been proposed by some process.
Lemma 11. Byzantine MPTC satisfies agreement.

Proof. Note that Phase 1 respects agreement by the correctness of the protocols in $\mathcal{P}$. Observe that honest processes can only decide during Phase 3 if at least $f + 1$ processes have decided the same value. This means that if an honest process decides $v$ during Phase 3 of MPTC at least one honest process has already decided $v$. Using similar arguments to those in Lemma 2, we can show that at any point during MPTC execution any two honest processes that decide must decide the same value.

Lemma 12. Byzantine MPTC satisfies unanimity.

Proof. Assume that all honest processes are initialized with the same value $v \in \mathcal{V}$. Since the BFT protocols in $\mathcal{P}$ are correct, honest processes getting into Phase 1 with the same value, $v$, can only compute outcomes $(D, v)$, $(M, v)$ and $(U, v)$. Otherwise it would mean that honest processes can change their proposals between rounds of the BFT protocol in question (e.g., via the influence of the faulty processes) and eventually decide a different value, which would violate unanimity of the BFT protocol. Therefore, at the end of Phase 1 all honest processes will end up with outcomes containing $v$. $v$ will be the most frequent value in any subset of $\mathcal{P}_f - f$ outcomes since $\mathcal{P}_f \geq 3f + 1$ and at least $2f + 1$ processes are correct. As a result Phases 2 and 3 will have all honest processes updating their outcomes and proposals to $v$. Consequently, all honest processes will eventually move to subsequent rounds with $v$ as their proposal and thus $v$ will be the only value that can be decided by some honest process.

C Implementation Details

C.1 MPTC protocol implementation

To implement MPTC, we need to decide on the following parameters:

- The choice of protocol set $\mathcal{P}$.
- The set of possible configurations $\mathcal{C}$, which specifies not only the protocol of each round but also its initialization.
- The configuration selection of functions $F_S$, $\forall S \in \mathcal{P}_S$, generated by the trusted dealer.
- The implementation of the split function.
- The implementation of GFS and combine used during the main protocol.

Our set of protocols, $\mathcal{P}$, contains only a single consensus protocol, an optimized version of single decree Paxos [18]. A round of single decree Paxos operates in 2 phases. In the first phase an active leader/proposer gets elected and in the second the active leader makes its proposal to the rest of the processes who may accept the proposal. If the proposal gets accepted by a majority of acceptors, it gets decided. Paxos tolerates $f$ crash failures using $2f + 1$ processes and thus $\mathcal{P}_f = 2f + 1$.

Similar to the implementation of Turtle Consensus [22], we avoid electing leaders in each round by using a parameterized version of single decree Paxos in which each round comes with a predetermined leader known to all active participants. Assuming an ordering of the processes executing the protocol, we can set the leader to be the process whose position in that order is equal to round number modulo $2f + 1$. Under normal execution conditions, the previous optimization yields the following benefits: 1) a decision within a single round-trip of communication since it directly executes Phase 2 of Paxos, and 2) performance unaffected by contention since the active leader is the only proposer.
For handling failures, our optimization assumes the weakest failure detector, \(\diamond W\), presented in [7]. If a leader failure is suspected by \(\diamond W\), then the suspecting processes will complete that round using \(M\) or \(U\) outcomes depending on whether they have received a proposal or not, respectively. We implement \(\diamond W\) like we did in Turtle Consensus, using timeouts with exponentially increasing timeout periods when processes are inaccurately suspected. This way we ensure that there will be enough rounds executed “concurrently” by sufficiently many processes which is critical for ensuring termination in our Paxos variant.

In the description above we did not specify how the states of processes running Paxos are turned into outcomes at the end of a Paxos round. The process executing as the active leader can either decide the value it proposes during Phase 2, say \(v\), or fail to do so due to either failing or suspecting a majority of acceptors. In the first case, it computes outcome \((D, v)\). In the second case and if the leader did not fail, it will timeout knowing that if a decision was made it must have been for its own proposal, thus computing outcome \((M, v)\), where \(v\) is its proposal in the beginning of the round. Similarly, a process running as an acceptor can end a round either having accepted the active leader’s proposal, thus computing outcome \((M, v)\), or timing out without having received any proposal, in which case it does not know anything about the decision progress. In this latter case, we need to ensure that if the acceptor becomes the next round’s active leader, it will make a proposal consistent with already accepted proposals. The way original Paxos achieves this, is through its phase 1. As a result our variant requires processes exiting the Paxos round without knowledge of the round’s decision state to retrieve this knowledge from the rest of the processes.

To avoid incurring another round of communication in our Paxos variant we piggyback this decision state retrieval onto Phase 2 of MPTC. Timed out processes can use the set of outcomes received to update their proposal. The update procedure is similar to the one used by processes that have computed outcome \((U, \ast)\). The main difference is that processes without knowledge about the outcome of the round send a special “unknown” outcome. These “unknown” outcomes are ignored by receiving processes unless all outcomes received during Phase 2 of MPTC are “unknown”, in which case no decision could have been made. In the latter case, the receiving process updates its outcome to \((U, v')\), where \(v'\) is the receiver’s proposal at the beginning of the round, and acts as in Phase 2, Step 2, Case 3. The above optimization ensures that if the predetermined leader decides value \(v\) at some round \(r\) and a failure prevents that decision to be learnt in \(r\), then all processes during Phase 2 of \(r\) will receive at least one \((M, v)\) outcome. The receiving processes will be forced to adopt \(v\) and thus future proposals can only be about \(v\).

Our set of configurations is simply \(C = \{(S, P) | S \in PS\}\) where \(P \in \mathcal{P}\) is the prior Paxos variant. We assume that for every participant set there is an ordering of the processes in it. This is easy to achieve using the processes unique identifiers, and it facilitates the selection of leader of each configuration without having to define additional initialization information in each configuration. Note that since all rounds execute the same correct consensus protocol, all possible interleavings in \(C\) are conforming.

Observe that in contrast to our prior work on Turtle Consensus we use the same protocol across configurations. In Turtle Consensus [22], different configurations used the same \(2f + 1\) set of processes. As a result, the adversary could try to track the current leader within that set of processes even if the leader changed across different configurations. Therefore, a competent adversary could eventually locate and force Turtle Consensus rounds to fail which can lead to very poor performance. For that reason, we kept switching between a leader-based (Paxos) and fully decentralized (Ben-Or) consensus protocols across configurations to prevent the adversary of exploiting the leader vulnerability. A side-effect of that approach, however,
was that by falling back to a less efficient protocol (Ben-Or) we only achieved sub-par performance compared to the graceful execution using only Paxos rounds. With MPTC we do not need to employ such tactics since the adversary now needs to scan through $|N| \gg f$ processes before it can identify the leader of our Paxos configuration.

The remaining parameters of our implementation are related to the cryptographic framework assumed by our protocol. While we did not implement this framework for our evaluation, we describe for completeness how we could do so in the following paragraphs. For the actual implementation that we evaluate in Section 5, we assume that all participant sets use the same unpredictable function given to all processes via a configuration file. This file simply defines a sequence of configurations, one for each round, that processes move to in a round-robin fashion, that is use the first configuration in round 0, the second in round 1, etc. Once the last configuration in the sequence is used, the processes loop back to the first one and continue from there. We emulate the restrictions that the cryptographic framework imposes on the adversary by assuming that only the processes involved in rounds $r$ and $r+1$ can learn $C_{r+1}$ and only after Phase 2 of round $r$ completes.

A potential implementation for the unpredictable functions is having $F_S = F$, $\forall S \in PS$, where $F$ is derived from the threshold coin-tossing scheme implementation based on Diffie-Hellman and presented in [5]. This approach hashes the input value, $r$, using a cryptographically secure hash function, modeled as a random oracle, and raises the result to a secret exponent. This exponent is shared among the processes using Shamir’s secret sharing [24]. Finally, the result is further hashed to obtain the value of $F(r)$ using another cryptographically secure hash function.

Function $split$ used by the dealer during initialization can be implemented using Shamir’s secret sharing as mentioned above. Function $GFS$ can be implemented by having each process $p$ hashing the input round number $r$ and raising it to its secret share of the exponent it received from the dealer. Finally, $combine$ for (byzantine) MPTC simply multiplies a set of $f+1$ distinct (valid) shares and hashes the result. Note that the $combine$ computation is slightly more complicated and a detailed version of its implementation can be found in [5]. Also note that one can alternatively use any non-interactive threshold-signature scheme with the property that there is only one valid signature per message, like the RSA-based scheme of Shoup [25]. We can then obtain the value of the function by hashing the resulting signature computed by $f+1$ signature shares on a message containing input $r$.

### C.2 MPTC-based state machine replication

In this section we describe in detail how we implemented SMRP using our MPTC implementation. Messages exchanged in this implementation are of the form:

$$(type, src, dst, (content-\text{-}attribute-1, content-\text{-}attribute-2,\ldots))$$

where $type \in \{\text{REQUEST, RESPONSE, DECISION, RECONFIGURATION}\}$, $src, dst \in \mathbb{N}$ are the ids of the communicating processes, and $content-\text{-}attribute-x$, for $x \in \mathbb{N}$ constitute the message payload. The different types of messages are as follows: A request message is sent by a client to a participant and may be forwarded between participants. It carries a deterministic operation to be executed by the service. A response message is sent by a replica to a participant which then forwards it to clients and contains the result of the execution of an operation issued via a request by the client. A decision message is sent by a participant to a replica, and carries a request along with its order of execution. Finally, reconfiguration
Moving Participants Turtle Consensus

messages are sent between participants and are used to update the configuration of the MPTC execution.

C.2.1 Clients

Clients are uniquely identified processes whose identities are independent from those of the participants and the replicas. They connect to at least \( f + 1 \) participants to which they send requests of the form:

\[(\text{client-id}, \text{request-number}, \text{command})\]

where \( \text{request-number} \) is a unique identifier for a particular request sent by this client and \( \text{command} \) is an application-specific description of a deterministic operation and of any arguments that operation requires. Note that each pair \((\text{client-id}, \text{request-number})\) uniquely identifies a request received by the state machine and will be used by participants and replicas to track requests that are new, under processing, or executed. Clients therefore maintain the following state:

- \( \text{cid} \in \mathbb{N} \), which is initialized with a unique value identifying the client.
- \( \text{rsn} \in \mathbb{N} \), which is initialized to 0, incremented each time the client issues a new request and uniquely identifies requests for a given client.
- \( \text{pending-requests} \subseteq \mathbb{N} \), which is an initially empty set that stores ids of requests the client has sent, but has not yet received response.

Our clients are modeled as state machines with the following transitions:

T1: **Precondition:** There is a command \( \text{cmd} \) that needs to be executed

**Action:** Send message \((\text{cid}, \text{pid}, \text{REQUEST}, (\text{cid}, \text{rsn}, \text{cmd}))\) to each participant \( \text{pid} \) that \( \text{cid} \) is connected to; add \( \text{rsn} \) to \( \text{pending-requests} \) and update \( \text{rsn} = \text{rsn} + 1 \).

T2: **Precondition:** Received message \((\text{pid}, \text{cid}, \text{RESPONSE}, (\text{rsn}', \text{result}))\)

**Action:** If \( \text{rsn}' \notin \text{pending-requests} \) ignore; otherwise remove \( \text{rsn}' \) from \( \text{pending-requests} \).

C.2.2 Participants

The participants receive requests from clients and are responsible for ordering these requests and send them for execution to the replicas. To achieve this they spawn MPTC instances, one for each request that needs to be ordered. Each instance has its own identifier and decided requests are ordered according to the identifiers of the MPTC instances that decided them. Only one participant set can be active at any point in time. The participants of this set are responsible for creating consensus instances. The active participant set may concurrently run different instances of MPTC to deal with multiple requests. As in the case of Turtle Consensus [22], we consider a global threshold \( W \) that limits the number of concurrent consensus instances that have not yet decided. Requests that arrive when \( W \) concurrent MPTC instances are running are queued and processed when some of the running instances complete.

Note that in most common SMRPs clients typically send their requests to replicas directly, which are responsible for instantiating consensus instances to these requests. By disabling communication between clients and replicas we prevent clients from launching DoS attacks on the replicas. This approach does however have the issue that when a decision is made and the corresponding request is executed, two hops of communication are needed for the
response to arrive at the clients. In addition, it burdens participants with implementing functionality beyond running consensus that is commonly offered by replicas, like keeping track of the values already decided, maintain state for ongoing instances and other aspects of the implementation, which makes it a less efficient approach. In this implementation, however, we are primarily interested in building an attack-tolerant SMRP and thus made the choice of communication pattern we described above.

Each participant is connected to all replicas and zero or more clients. Apart from the state required to run MPTC described in Section 3, each participant additionally maintains the following state:

- \( \text{pid} \in N \), initialized with the identifier of the participant.
- \( \text{next-instance} \in N \), initialized to 0, stores the id of the next instance that will be created.
- \( \text{configuration} \in R \), initialized to \( C_0 \) provided by the dealer, stores the configuration that \( \text{pid} \) considers active.
- \( \text{requests} \subseteq N^2 \times \text{Ops} \), which is an initially empty set that stores requests that have been received but not yet decided. \( \text{Ops} \) denotes the space of commands and it is application-specific.
- \( \text{instances} \subseteq N^3 \), which is an initially empty set that stores for each running instance the id of the instance as well as the request identifier \( (\text{cid}', \text{rsn}) \); the request identifier is \( \text{pid} \)'s input proposal for that MPTC instance.
- \( \text{rstate} \in \{\text{TRUE}, \text{FALSE}\} \), initialized to \text{FALSE}, indicates whether the participant is currently under reconfiguration.
- \( \text{responses} \subseteq N^3 \), which is an initially empty set that stores mappings of requests and processes from which they were received. It is primarily used to forward responses back to the processes that sent or relayed requests.

Let \( \text{configuration.participants} \) denote the participant set of \( \text{configuration} \) and let \( \text{configuration.round} \) denote the round in which \( \text{configuration} \) is used. Aside from the transitions determined by the MPTC instances a participant may be running, it additionally performs the following transitions:

**T1:** **Precondition:** Received message \( \langle \text{REQUEST}, \text{pid}', \text{pid}, (\text{cid}, \text{rsn}, \text{cmd}) \rangle \) where \( \text{pid}' \) is either a client or another participant, \( \text{rstate} = \text{FALSE}, \text{pid} \notin \text{configuration.participants} \) and \( (\text{pid}', \text{cid}, \text{rsn}) \notin \text{responses} \)

**Action:** Add \( (\text{pid}', \text{cid}, \text{rsn}) \) in \( \text{responses} \) and send \( \langle \text{REQUEST}, \text{pid}, (\text{cid}, \text{rsn}, \text{cmd}) \rangle \) to each \( p \in \text{configuration.participants} \).

**T2:** **Precondition:** Received message \( \langle \text{REQUEST}, \text{pid}', \text{pid}, (\text{cid}, \text{rsn}, \text{cmd}) \rangle \) where \( \text{pid}' \) is either a client or another participant, \( \text{rstate} = \text{FALSE}, \text{pid} \in \text{configuration.participants}, \exists (\text{cid}, \text{rsn}, *) \in \text{requests} \) and \( (\text{pid}', \text{cid}, \text{rsn}) \notin \text{responses} \)

**Action:** Add \( (\text{cid}, \text{rsn}, \text{cmd}) \) in \( \text{requests} \) and \( (\text{pid}', \text{cid}, \text{rsn}) \) in \( \text{responses} \). Send \( \langle \text{REQUEST}, \text{pid}, (\text{cid}, \text{rsn}, \text{cmd}) \rangle \) to each \( p \in \text{configuration.participants} \setminus \{\text{pid}\} \). If \( W > |\text{instances}| \) then create an MPTC instance with instance id, \text{next-instance} and proposal \( (\text{cid}, \text{rsn}) \). Then add \( \langle\text{next-instance}, \text{cid}, \text{rsn}\rangle \) to \( \text{instances} \) and update \( \text{next-instance} = \text{next-instance} + 1 \). Finally, run round \( \text{configuration.round} \) of the new MPTC instance.

**T3:** **Precondition:** Received message \( \langle \text{REQUEST}, \text{pid}', \text{pid}, (\text{cid}, \text{rsn}, \text{cmd}) \rangle \) where \( \text{pid}' \) is either a client or another participant, \( \text{rstate} = \text{FALSE}, \text{pid} \in \text{configuration.participants}, \exists (\text{cid}, \text{rsn}, *) \in \text{requests} \) and \( (\text{pid}', \text{cid}, \text{rsn}) \notin \text{responses} \)

**Action:** Add \( (\text{pid}', \text{cid}, \text{rsn}) \) in \( \text{responses} \).
T4: **Precondition:** Received message \(<\text{REQUEST}, \text{pid}', \text{pid}, (\text{cid}, \text{rsn}, \text{cmd})>\) where \(\text{pid}'\) is either a client or another participant, and \(\text{rstate} = \text{TRUE}\)
**Action:** If \(\nexists (\text{cid}, \text{rsn}, *) \in \text{requests}\) add \((\text{cid}, \text{rsn}, \text{cmd})\) in \(\text{requests}\). If \((\text{pid}', \text{cid}, \text{rsn}) \notin \text{responses}\) add \((\text{pid}', \text{cid}, \text{rsn})\) in \(\text{responses}\).

T5: **Precondition:** Received message \(<\text{RESPONSE}, \text{rid}, \text{pid}, (\text{cid}, \text{rsn}, \text{result})>\) from any replica or participant \(\text{rid}\) and \(\exists (*, \text{cid}, \text{rsn}) \in \text{responses}\)
**Action:** For each \((\text{pid}', \text{cid}, \text{rsn}) \in \text{responses}\) send \(<\text{RESPONSE}, \text{pid}, \text{pid}', (\text{cid}, \text{rsn}, \text{result})>\) and remove \((\text{pid}', \text{cid}, \text{rsn})\) from \(\text{responses}\).

T6: **Precondition:** On round completion of some instance \((\text{instance}, \text{cid}, \text{rsn}) \in \text{instances}\) with outcome \((\text{D}, (\text{cid}, \text{rsn}))\) and \((\text{cid}, \text{rsn}, \text{cmd}) \in \text{requests}\)
**Action:** Send \(<\text{DECISION}, \text{pid}, \text{rid}, (\text{instance}, \text{cid}, \text{rsn}, \text{cmd}, \text{configuration})>\) to each \(\text{rid} \in \mathcal{R}\). Then remove \((\text{instance}, \text{cid}, \text{rsn})\) from \(\text{instances}\) and \((\text{cid}, \text{rsn}, \text{cmd})\) from \(\text{requests}\).

T7: **Precondition:** \(|\text{instances}| = 0, \exists (\text{cid}, \text{rsn}, \text{cmd}) \in \text{requests}, \text{rstate} = \text{FALSE}\) and \(\text{pid} \in \text{configuration.participants}\)
**Action:** Create an MPTC instance with instance id, \(\text{next-instance}\) and proposal \((\text{cid}, \text{rsn})\). Then add \((\text{next-instance}, \text{cid}, \text{rsn})\) to \(\text{instances}\) and update \(\text{next-instance} = \text{next-instance} + 1\). Finally, run round \(\text{configuration.round}\) of the new MPTC instance.

T8: **Precondition:** On round completion of some instance \((\text{instance}, \text{cid}, \text{rsn}) \in \text{instances}\) with outcome \((\text{M}, (\text{cid}, \text{rsn}))\) or \((\text{U}, (\text{cid}, \text{rsn}))\) and there are still instances that have not completed \(\text{configuration.round}\)
**Action:** Update \(\text{rstate} = \text{TRUE}\).

T9: **Precondition:** On round completion of some instance \((\text{instance}, \text{cid}, \text{rsn}) \in \text{instances}\) with outcome \((\text{M}, (\text{cid}, \text{rsn}))\) or \((\text{U}, (\text{cid}, \text{rsn}))\), next configuration \(\text{cfg}\) and there no more instances that have not completed \(\text{configuration.round}\)
**Action:** Update \(\text{rstate} = \text{FALSE}\) and \(\text{configuration} = \text{cfg}\). Then send \(<\text{RECONFIGURATION}, \text{pid}, \text{p}, (\text{instances}, \text{requests}, \text{configuration}, \text{instance})>\) to each \(\text{p} \in \text{configuration}\). Finally, update \(\text{instances} = \emptyset\) and \(\text{requests} = \emptyset\).

T10: **Precondition:** Received messages \(<\text{RECONFIGURATION}, \text{p}, \text{pid}, (\text{instances}_p, \text{requests}_p, \text{configuration}', \text{instance}_p)>\) from a set of participants \(\mathcal{Q}, |\mathcal{Q}| = f + 1\) such that \(\text{configuration}', \text{round} > \text{configuration.round}\)
**Action:** Update \(\text{configuration} = \text{configuration}'\), \(\text{next-instance} = \text{max}(\text{instance}_p, \text{for } \text{p} \in \mathcal{Q})\) and \(\text{requests} = \text{requests} \cup (\cup_{\text{p} \in \mathcal{Q}} \text{requests}_p)\). For each \(\text{p} \in \mathcal{Q}\) and \((\text{id}, \text{rsn}, \text{cmd}) \in \text{requests}_p\) add \((\text{p}, \text{cid}, \text{rsn})\) in \(\text{responses}\). For each \((\text{instance}', \text{cid}, \text{rsn}) \in \cup_{\text{p} \in \mathcal{Q}} \text{instances}_p\) use Phase 3 of MPTC to update the proposal of \(\text{instance}'\) and create an MPTC instance \((\text{instance}', \text{cid}', \text{rsn}')\) where \((\text{cid}', \text{rsn}')\) is the updated proposal. Finally, add \((\text{instance}', \text{cid}', \text{rsn}')\) to \(\text{instances}\).

Like in Turtle Consensus, MPTC is lazily instantiated for each slot. MPTC messages carry instance identifiers so incoming protocol messages are properly processed by the correct instance. If an instance has not yet been created, messages for that instance are queued and processed when it is created. Transition T2 ensures that if a request creates an instance to one of the active participants then unless the receiving participant fails, the remaining correct processes in the active participant set will also receive and create an instance for that request. Transition T7 ensures that if a proposed request does not get decided in one of the running instances, a new instance will be created for that proposal.
C.2.3 Replicas

The sets of replicas and participants are disjoint and $|\mathcal{R}| > f$ so at least one replica is always correct. Each replica in $\mathcal{R}$ is run by a different process and maintains a copy of state of the service implemented by the SMRP. Replicas are responsible for executing the commands issued through clients’ requests in the order established by the participants. This order is determined by the instance id’s associated with each decision received. As in [22], we model this ordering as a sequence of numbered slots with the first slot numbered as 0. All replicas are initialized in the same state and since all commands are deterministic, executing commands in order at all replicas ensures that they all end up in the same state. To provide this functionality, replicas maintain the following state additional to the one related with the service SMRP implements:

- $ri \in \mathcal{R}$, initialized with the identifier of the replica.
- $execution-slot \in \mathbb{N}$, which is initialized 0 and maintains the position in the ordering of commands that should be executed next.
- $decisions \subseteq \mathbb{N}^3 \times Ops \times \mathcal{C}$, which is an initially empty set of mappings between a slot and a request as well as the configuration under which the request was decided.
- $completed \subseteq \mathbb{N}^3$, which is an initially empty set that stores mappings between slots and ids of requests that were executed in those slots.

In addition, to those transitions defined by the deterministic state machine implemented by the SMRP, replicas perform the following transitions:

T1: **Precondition:** Received message $\langle DECISION, pid, rid, (instance, cid, rsn, cmd, C) \rangle$ where $C \in \mathcal{C}, \not\exists (instance, \ast, \ast, \ast, \ast) \in decisions$, and $execution-slot \leq instance$

**Action:** Add $(instance, cid, rsn, cmd, C)$ in decisions.

T2: **Precondition:** $\exists (instance, cid, rsn, cmd, C) \in decisions$, and $execution-slot = instance$, and $\exists (*, cid, rsn) \in completed$

**Action:** Execute $cmd$ and let $result$ be the outcome of the operation. Send $\langle RESPONSE, rid, pid, (cid, rsn, result) \rangle$ to each $pid$ in the participant set specified by $C$. Add $(instance, cid, rsn)$ to completed and remove $(instance, cid, rsn, cmd, C)$ from decisions. Update $execution-slot = execution-slot + 1$.

T3: **Precondition:** $\exists (instance, cid, rsn, cmd, C) \in decisions$, and $execution-slot = instance$, and $\exists (*, cid, rsn) \in completed$

**Action:** Add $(instance, cid, rsn)$ to completed and remove $(instance, cid, rsn, cmd, C)$ from decisions. Update $execution-slot = execution-slot + 1$.

In other words, replicas execute commands one slot at a time as soon as they become available. Decisions already received or executed are ignored or treated as no-op. Note that there is still the issue of garbage collecting executed requests in $completed$ that are no longer needed. In fact, $completed$ can get arbitrarily large the longer the system operates. One solution would be to only track for each client the largest request identifier used such that future decisions about an already executed request number for a given client can be ignored.