Distributed Transactions: Dissecting the Nightmare

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Abstract

Many distributed storage systems are transactional and a lot of work has been devoted to optimizing their performance, especially the performance of read-only transactions that are considered the most frequent in practice. Yet, the results obtained so far are rather disappointing, and some of the design decisions seem contrived. This paper contributes to explaining this state of affairs by proving intrinsic limitations of transactional storage systems, even those that need not ensure strong consistency but only causality.

We first consider general storage systems where some transactions are read-only and some also involve write operations. We show that even read-only transactions cannot be "fast": their operations cannot be executed within one round-trip message exchange between a client seeking an object and the server storing it. We then consider systems (as sometimes implemented today) where all transactions are read-only, i.e., updates are performed as individual operations outside transactions. In this case, read-only transactions can indeed be "fast", but we prove that they need to be "visible". They induce inherent updates on the servers, which in turn impact their overall performance.

1 Introduction

Transactional distributed storage systems have proliferated in the last decade: Amazon's Dynamo [1], Facebook's Cassandra [2], Linkedin's Espresso [3], Google's Megastore [4], Walter [5] and Lynx [6] are seminal examples, to name a few. A lot of effort has been devoted to optimizing their performance for their success heavily relies on their ability to execute transactions in a fast manner [7]. Given the difficulty of the task, two major "strategic" decisions have been made. The first is to prioritize *read-only transactions*, which allow clients to read multiple items at once from a consistent view of the data-store. Because many workloads are read-dominated, optimizing the performance of read-only transactions has been considered of primary importance. The second is the departure from strong consistency models [8, 9] towards weaker ones [10, 11, 12, 13, 14, 15]. Among such weaker consistency models, *causal consistency* has garnered a lot of attention for it avoids heavy synchronization inherent to strong consistency, can be implemented in an always-available fashion in geo-replicated settings (i.e., despite partitions), while providing sufficient semantics for many applications [16, 17, 18, 19, 20, 21, 22]. Yet, even the performance of highly optimized state-of-the-art causally consistent transactional storage systems has revealed disappointing. In fact, the benefits and implications of many designs are unclear, and their overheads with respect to systems that provide no consistency are not well understood.

To illustrate this situation, we report here on two state-of-the-art designs. The first implements what we call "fast" read-only transactions. They complete in one round of interaction between a client seeking to read the value of an object and the server storing it. This design is implemented by

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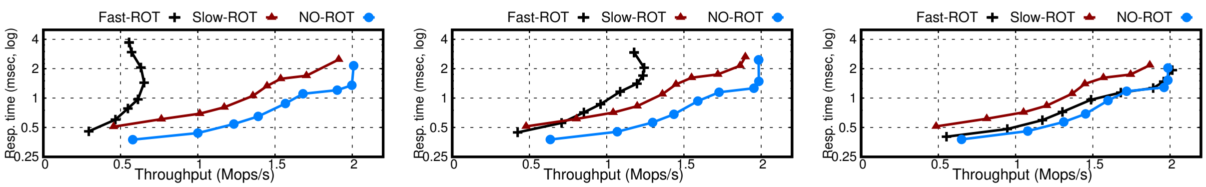
the recent COPS-SNOW [12] system, which however makes the assumption that write operations are supported only outside the scope of a transaction.[1] The second design implements "slow" read-only transactions, that require two communication rounds to complete. In particular, we consider the design of Cure [21], which supports generic read-write transactions. We compare these two systems with three read-dominated workloads corresponding to 0.999, 0.99 and 0.95 read-write ratios, where clients perform read-only transactions and single-object write operations in closed loop.[2] Figure 1 reports on the average latency of read-only transactions for the two designs ("fast" and "slow") as a function of the delivered throughput, and compares them with those achieved by a system that guarantees no consistency ("no"). The plots depict two results. First, the slow case results in a higher latency with respect to a design with only one round and no consistency. This raises the question whether it is possible to preserve the rich semantics of generic read-write transactions and implement read-only transactions with a single communication round. Second, the performance achieved by the fast read-only transactions are worse than the ones achieved by the slow ones, both in latency and throughput, even for read-write ratios as low as 5%. This is unexpected.



(a) 5% writes (b) 1% writes (c) 0.1% writes

Figure 1: Performance of "fast", "slow" transactions, and no transaction guarantees

In this paper, we investigate these aspects from a theoretical perspective with the aim of identifying possible and impossible causal consistency designs in order to ultimately understand their implications. We prove two impossibility results.

- First, we prove that no causally consistent system can support generic transactions and implement fast read-only transactions. This result unveils a fundamental trade-off between semantics (support for generic transactions) and performance (latency of read-only transactions).
- Second, we prove that fast read-only transactions must be "visible", i.e., their execution updates the states of the involved servers. The resulting overhead increases resource utilization, which sheds light on the inherent overhead of fast read-only transactions and explains the surprising result discussed before.

The main idea behind our first impossibility result is the following. One round-trip message exchange disallows multiple servers to synchronize their responses to a client. Servers need to be conservative and return possibly stale values to the client in order to preserve causality, with the risk of jeopardizing progress. Servers have no choice but communicate outside read-only transactions (i.e., helping each other) to make progress on the freshness of values. We show that such message exchange can cause an infinite loop and delay fresh values forever. The intuition behind our second

[1]Under this assumption, single-object write and a transaction that only writes to one object are equivalent.

[2]We implemented these in the same C++ code-base using Google Protobuf library for communication. We run the workload on a 10Gbps Ethernet network using 64 AMD Opteron 6212 machines with 8 physical cores (16 hardware threads) and 64 GB of RAM (where 32 machines host client processes, and 32 host server processes).

result is different. We show that a fast read-only transaction has to "write" to some server for otherwise, a server can miss the information that a stale value has been returned for some object by the transaction (which reads multiple objects), and return a fresh value for some other, violating causal consistency.

At the heart of our results lies essentially a fundamental trade-off between causality and (eventual) freshness of values.[3] Understanding this trade-off is key to paving the path towards a new generation of transactional storage systems. Indeed, the relevance of our results goes beyond the scope of causal consistency. They apply to any consistency model stronger than causal consistency, e.g., linearizability [8, 9] and strict serializability [23, 24], and are relevant also for systems that implement hybrid consistency models that include causal consistency, e.g., Gemini [25] and Indigo [26].

The rest of this paper is organized as follows. Section 2 presents our general model and definitions. Section 3 presents the impossibility of fast read-only transactions. Section 4 presents the impossibility of fast invisible read-only transactions (in the restricted model). Section 5 discusses related work. Section 6 discusses how to circumvent our impossibility results. For space limitation, we defer the details of the proofs of our impossibility results to the appendix.

2 Model and Definitions

2.1 Model

We assume an arbitrarily large number of *clients* $C_1, C_2, C_3, \ldots$ (sometimes also denoted by C), and at least two *servers* P_X, P_Y. Clients and servers interact by exchanging messages. We consider an *asynchronous* system where the delay on message transmission is finite but arbitrarily large, and there is no global clock accessible to any process. Communication channels do not lose, modify, inject, or duplicate messages, but messages could be reordered.

A *storage* is a finite set of objects. Clients read and/or write objects in the storage via *transactions*. Any transaction T consists of a read set R_T and a write set W_T on an arbitrary number of objects (R_T or W_T could be empty). We denote T by (R_T, W_T). We say that a client *starts* a transaction when the client *requests* the transaction from the storage. Any client which requests transaction T returns a value for each read in R_T and *ok* for each write in W_T. We say that a client *ends* a transaction when the client returns from the transaction. Every transaction ends.

The storage is implemented by servers. For simplicity of presentation, we assume that each server stores a different set of objects and the set is disjoint among servers. (We show in the appendix how our results apply to the non-disjoint case.) Every server receiving a request from a client responds. A server's response without any client request is not considered, and no server receives requests for objects not stored on that server. Naturally, a server that does not store an object stores no information on values written to that object. Clients do not buffer the value of an object to be read; instead a server returns one and only one value which has been written to the object in question.

2.2 Causality

We consider a storage that ensures *causality* in the classical sense of [27], which we first recall and adapt to a transactional context.

[3]This trade-off is different from the traditional one in distributed computing between ensuring linearizability (i.e., finding a linearization point) and ensuring wait-freedom, both rather strong properties.

The *local history* of client C_i, denoted L_i, is a sequence of start and end events. We assume, w.l.o.g., that any client starts a new transaction after the client has ended all previous transactions, i.e., clients are sequential. Hence any local history L_i can be viewed as a sequence of transactions. We denote by $r(x)v$ a read on object x which returns v, by $r(x)*$ a read on object x for an unknown return value (with symbol $*$ as a place-holder), and by $w(x)v$ a write of v to object x. For simplicity, we assume that every value written is unique. (Our results hold even when the same values can be written.) Definition 1 captures the program-order and read-from causality relation [27].

Definition 1 (Causality [27])**.** Given local histories $L_1, L_2, L_3 \ldots$, for any $\alpha = a(x_\alpha)v_\alpha, \beta = b(x_\beta)v_\beta$ where $a, b \in \{r, w\}$, we say that α causally precedes β, which we denote by $\alpha \rightsquigarrow \beta$, if (1) $\exists i$ such that α is before β in L_i; or (2) $\exists v, x$ such that $\alpha = w(x)v$ and $\beta = r(x)v$; or (3) $\exists \gamma$ such that $\alpha \rightsquigarrow \gamma$ and $\gamma \rightsquigarrow \beta$.

Our definition of causally consistent transactions follows closely the original definition of [27]. We only slightly extend the classical definition of causal serialization in [27] to cover transactions. Assume that each object is initialized with a special symbol $\perp$. (Thus a read can be $r(x)\perp$.)

Definition 2 (Transactional causal serialization)**.** Given local histories $H = L_1, L_2, L_3, \ldots$, we say that client C_i's history can be causally serialized if we can totally order all transactions that contain a write in H and all transactions in L_i, such that (1) for any read $r = r(x)v$ on object x which returns a non-$\perp$ value v, the last write $w(x)v_w$ on x which precedes the transaction that contains r satisfies $v_w = v$; (2) for any read $r = r(x)\perp$ on object x, no write on x precedes the transaction that contains r; (3) for any α, β such that $\alpha \rightsquigarrow \beta$, the transaction that contains α is ordered before the transaction that contains β.

Definition 3 (Causally consistent transactional causal storage)**.** We say that storage *cc* is causally consistent if for any execution of clients with *cc*, each client's local history can be causally serialized.

2.3 Progress

Progress is necessary to make any storage useful; otherwise, an implementation which always returns $\perp$ or values written by the same client can trivially satisfy causal consistency. To ensure progress, we require any value written to be eventually *visible*. While rather weak, this definition is strong enough for our impossibility results, which apply to stronger definitions. We formally define progress in Definition 4 below. Existing implementations of causal consistency [16, 28, 17, 19, 21, 22, 29] indeed used the terminology of visible writes/updates/values and implicitly included progress as a property of their causally consistent systems, yet there has been no formal definition for progress.[4]

Definition 4 (Progress)**.** A (causally consistent) storage guarantees progress if, for any write $w = w(x)v$, v is eventually visible: there exists finite time $\tau_{x,v}$ such that any read $r(x)v_{new}$ which starts at time $t \geq \tau_{x,v}$, satisfies $v_{new} = v$ or $w(x)v_{new}$ returns no earlier than w starts.[5]

3 The Impossibility of Fast Transactions

In this section, we present and prove our first theoretical result, Theorem 1. We first define formally the notion of *fast* transactions. In short, a fast transaction is one of which each operation executes in (at most) one communication round between a client and a server (Definition 5 below).

[4] Bailis et al. [18] defined eventual consistency in a similar way to progress here; however they considered progress only in the situation where all writes can stop.

[5] The accurate time is used for the ease of presentation for definitions and proofs and not accessible to any process.

Definition 5 (Fast transaction)**.** We say that transaction T is fast if for any client C and C's invocation I of T, there is an execution where I ends and during I, for any server P:

- C sends at most one message to P and receives at most one message from P;
- If C sends a message to P, then after the reception of that message, any message which P sends to a server is delayed and P receives no message from any server until I ends.

Definition 5 excludes implementations where a server waits for the reception of messages from another server (whether the server is one which C sends a message to or not) to reply to a client. Definition 5 allows parallel transactions.

3.1 Result

Theorem 1 says that it is impossible to implement fast transactions (even if just read-only ones are fast).

Theorem 1. *If a causally consistent transactional storage provides transactions that can read and/or write multiple objects, then no implementation provides fast read-only transactions.*

The intuition behind Theorem 1 is the following. Consider a server P_X that stores object X and a server P_Y that stores object Y. If there is a risk of violating causality for P_Y where P_X could return an old value, then P_Y must also return an old value to the same transaction. In order to guarantee progress, extra communication is needed, which could further delay P_Y from returning a new value, in turn, creating a risk of violating causality for P_X. In fact, P_X and P_Y could take turns creating causality violation risks for each other, and preventing each other from returning new values forever, jeopardizing thereby progress. For space limitation, we just sketch below our proof of Theorem 1. (The full proof is deferred to the appendix.)

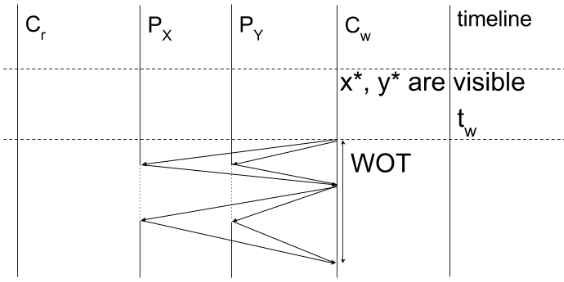
3.2 Proof overview

The proof of Theorem 1 is by construction of a contradictory execution E_{imp} which, to satisfy causality, contains an infinite number of messages the reception of which is necessary for some value to be visible (violating progress). As illustrated in Figure 2a, some non-$\perp$ values of X and Y have been visible in E_{imp}; then client C_w issues transaction $WOT = (w(X)x, w(Y)y)$ which starts at time t_w; since t_w, WOT is the only executing transaction. We make no assumption on the execution of WOT.

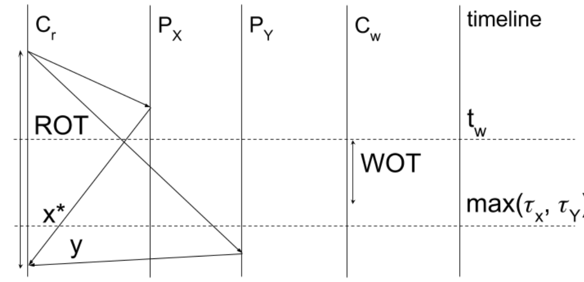
We show an infinite number of messages by induction on the number k of messages: no matter how many k messages have been sent and received, an additional message is necessary for x and y to be visible. Let $m_0, m_1, \ldots, m_{k-1}, m_k$ be the sequence of messages for case k. We show that in E_{imp}, except for m_0 and m_1, every message is sent after the previous message has been received. At the end of the induction, we conclude that in E_{imp}, these messages delay both x and y from being visible. As every message is sent after previous messages have been received, the delay accumulates and thus violates Definition 4. We sketch below the proof of the base case and the inductive step.

3.3 Base case

We first define some terminology to unify the description of communication between P_X and P_Y no matter whether the communication is via some third server or not: we say that P_X (P_Y) sends a message which *precedes* some message that arrives at P_Y (P_X), in the sense defined below (Definition 6). Thus the case where P_X sends message m to server S and S forwards m to P_Y is covered.



(a) Construction of E_{imp}



(b) Contradictory execution for the existence of the first two messages

Figure 2: Illustration of E_{imp} and the base case

Definition 6. Message m_1 *precedes* message m_2 if (1) $m_1 = m_2$, or (2) a process sends m_2 after it receives m_1 or (3) there exists message m such that m_1 precedes m and m precedes m_2.

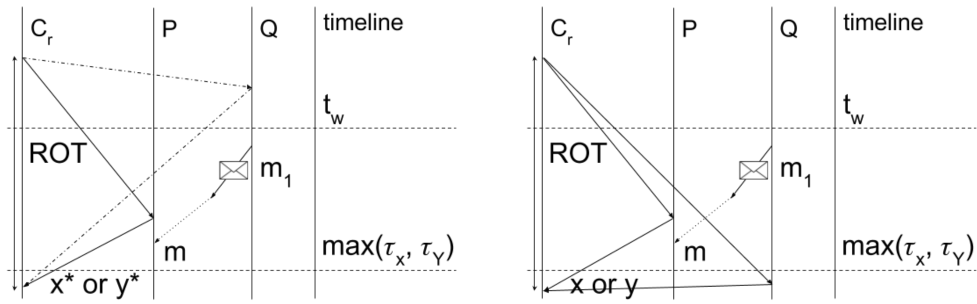
In the base case where $k = 1$, we show that after t_w, each server sends a message that precedes some message which arrives at the other server, by contradiction. By symmetry, suppose that P_X sends no message that precedes any message which arrives at P_Y. Then we add a read-only transaction ROT to E_{imp}, illustrated in Figure 2b: to P_X, the request of ROT is earlier than that of WOT, while to P_Y, the request of ROT is (much) later and is actually after x and y are eventually visible. By fast read-only transactions and our assumption for contradiction, after t_w, there can be no communication between P_X and P_Y before P_Y's response. As a result, ROT returns (x^*, y) for some $x^* \neq x$. Lemma 1 (of which the proof is also deferred to the appendix) depicts the very fact that such returned value violates causal consistency. A contradiction. By symmetry, we conclude that both P_X and P_Y have to send at least one message after t_w. These two messages are m_0 and m_1. Let $\{P, Q\} = \{P_X, P_Y\}$. Clearly, one server P between P_X and P_Y sends its message earlier than the other server Q. We let m_0 be the message sent by P and m_1, the other message.

Lemma 1. *In E_{imp}, no write (including writes in a transaction) occurs other than WOT since t_w. If some client C_r requests ROT, then ROT returns x if and only if ROT returns y.*

3.4 Inductive step

From case $k = 1$ to case $k = 2$, we show that another message m_2 is necessary (for the value written at Q to be visible). Let m be the first message which P receives and m_1 precedes. We argue by contradiction. Suppose that after the reception of m, P sends no message that precedes any message which arrives at Q. If the request of ROT comes at P after P sends m_0 and before P receives m, then by Lemma 1, P must return some $x^* \neq x$ or some $y^* \neq y$, considering the possibility that the request of ROT could come at Q before t_w, illustrated in Figure 3a. Now the request of ROT actually comes (much) later at Q (after the value written at Q is visible), illustrated in Figure 3b. By fast read-only transactions and our assumption for contradiction, after Q receives all messages preceded by m_0, there can be no communication between P and Q before Q's response. As a result, Q returns x or y and ROT returns (x^*, y) or (x, y^*), violating Lemma 1. A contradiction.

We then conclude that P must send m_2 after the reception of m, which is no earlier than the reception of m_1. For case $k = 3$, we can similarly show that Q must send another message m_3 (for the value written at P to be visible). In this way, we add one message in each step of the induction, while P_X and P_Y take turns in sending messages necessary for x and y to be visible. As



(a) Return of old values

(b) Contradictory execution

Figure 3: Existence of more messages after WOT

shown by induction, the total number of messages essentially grows to infinity. This completes our construction of E_{imp} as well as our proof sketch of Theorem 1.

4 The Impossibility of Fast Invisible Transactions

As we pointed out in the introduction, some systems considered a restricted model where all transactions are read-only and write operations are supported only outside the scope of a transaction. This restricted model also circumvents the impossibility result of Theorem 1. In this model, we present our second theoretical result, Theorem 2, stating that fast read-only transactions (while indeed possible) need to be visible (need to actually write). We first formally define the notion of (in)visible transactions in Definition 7 below.

Definition 7 (Invisible transactions)**.** We say that transaction T is invisible if for any client C and C's invocation I of T, any execution E (until I) can be continued arbitrarily but still there exists some execution E^- without I that is the same as E except for the message exchange with C (during the time period of I).

4.1 Result

Theorem 2 shows that it is impossible to implement fast invisible transactions (even if all transactions are read-only).

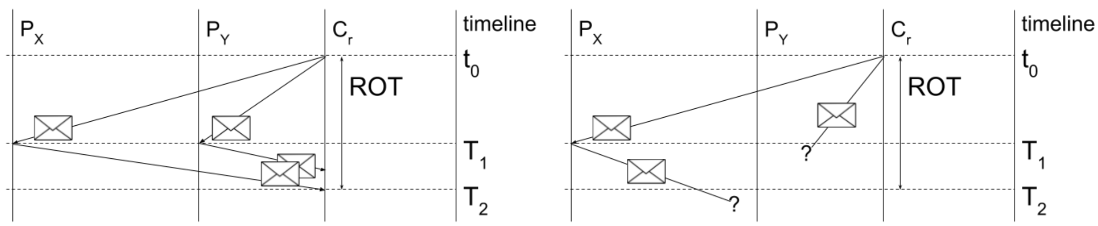
Theorem 2. *If a causally consistent transactional storage provides fast read-only transactions, then no implementation provides invisible read-only transactions.*

The intuition of Theorem 2 is the following. In an asynchronous system, any read-only transaction T can read an old value and a new value from different servers, and thus the communication that carries T is necessary to prevent T from returning a mix of old and new values. For space limitation, below we sketch our proof of Theorem 2. (The full proof is deferred to the appendix.)

4.2 Crucial executions

To prove Theorem 2, we consider any execution E_1 where some client C_r (which has not requested any operation before) starts transaction $ROT = (r(X)*, r(Y)*)$ at the time t_0. In E_1, before t_0, some values of X and Y have been visible. We continue E_1 with some client C executing $w(X)x$ and $w(Y)y$ (which establishes $w(X)x \rightsquigarrow w(Y)y$).

Our proof is by contradiction. Suppose that transaction ROT is invisible. Then no matter how E_1 is scheduled, there exists some execution E_2 such that E_2 is the same as E_1 except that (1) C_r does not invoke ROT, and (2) the message exchange with C_r during the time period of ROT is different. Below we first schedule E_1 and then construct another execution $E_{1,2}$. We later show $E_{1,2}$ violates causal consistency. By fast read-only transactions, we can schedule messages such that the message which C_1 sends during ROT arrives at P_X and P_Y respectively at the same time. Let T_1 denote this time instant and let T_2 be the time when ROT eventually ends, illustrated in Figure 4a. During $[T_1, T_2]$, P_X and P_Y receive no message but still respond to C_r. After T_2, the two writes of C occur, while C_r does no operation. All delayed messages eventually arrive before y can be visible. In E_1, y is visible after some time τ_y.



(a) Message schedule of E_1 (b) Message schedule of $E_{1,2}$

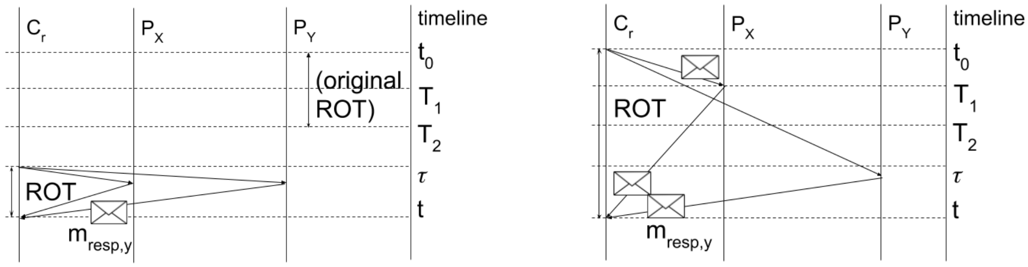
Figure 4: Construction and extension of E_1

Next we construct execution $E_{1,2}$ that is indistinguishable from E_1 to P_X and from E_2 to P_Y. The start of $E_{1,2}$ is the same as E_1 (as well as E_2) until t_0. At t_0, C_r still invokes ROT. As illustrated in Figure 4b, P_X receives the same message from C_r and sends the same message to C_r at the same time as in E_1; C_r sends the same message to P_Y at the same time as in E_1, the reception of which is however delayed by a finite but unbounded amount of time. In addition, during $[T_1, T_2]$, P_X and P_Y receive no message as in E_1 (as well as E_2). Thus by T_2, P_X is unable to distinguish between E_1 and $E_{1,2}$ while P_Y is unable to distinguish between E_2 and $E_{1,2}$. According to our assumption for contradiction, $E_{1,2} = E_1 = E_2$ except for the communication with C_r by T_2.

4.3 Proof overview

We continue our proof of Theorem 2 (by contradiction). Based on the executions constructed above, we extend E_2 and $E_{1,2}$ after τ_y. As illustrated in Figure 5, we let C_r start ROT immediately after τ_y in E_2. In both E_2 and $E_{1,2}$, by fast read-only transactions, we schedule the message sent from C_r to P_Y during C_r's ROT to arrive at the same time after τ_y, and $\exists t$ such that during $[\tau_y, t]$, P_Y receives no message but still responds to C_r. By t, P_Y is unable to distinguish between E_2 and $E_{1,2}$.

We now compute the return value of ROT in $E_{1,2}$. By progress, in E_2, P_Y returns y, and then by indistinguishability, in $E_{1,2}$, P_Y also returns y. Since in $E_{1,2}$, P_X returns some value $x^* \neq x$ (as $w(X)x$ starts after T_2), the return value of ROT in $E_{1,2}$ is (x^*, y). According to our assumption, $E_{1,2}$ satisfies causal consistency. By Definition 3, we can totally order all C_r's operations and all write operations in $E_{1,2}$ such that the last preceding writes of X and Y before C_r's ROT are $w(X)x^*$ and $w(Y)y$ respectively. This leads $w(X)x^*$ to be ordered after $w(X)x$. However, if we extend $E_{1,2}$ so that C_r invokes $ROT_1 = (r(X)*, r(Y)*)$ after x and y are visible, then ROT_1 returns value (x, y) and if we do total ordering of $E_{1,2}$ again, then the last preceding write of X before ROT_1 must be $w(X)x$, contradictory to the ordering between $w(X)x^*$ and $w(X)x$. Therefore, we conclude that $E_{1,2}$ violates causal consistency, which completes our proof sketch of Theorem 2.



(a) Extension of E_2

(b) Extension of $E_{1,2}$.

Figure 5: Extension of crucial executions

5 Related Work

5.1 Causal consistency

Ahamad et al. [27] were the first to propose *causal consistency* for a memory accessed by read/write operations. Bouajjani et al. [30] formalized the verification of causal consistency. A large number of systems [16, 28, 31, 17, 19, 22] implemented transactional causal consistency, although none formalized the concept for generic transactions. Akkoorath et al. [21] extended causal consistency to transactions by defining atomicity for writes and causally consistent snapshots for reads within the same transaction. Mehdi et al. [22] introduced *observable causal consistency* in the sense that each client observes a monotonically non-decreasing set of writes. Neither of the two definitions follows a formalization close to the original definition of [27].

5.2 Causal read-only transactions

Most implementations do not provide fast (read-only) transactions. COPS [16] and Eiger [17] provide a two-round protocol for read-only transactions. Read-only transactions in Orbe [31], GentleRain [19], Cure [21] and Occult [22] can induce more than one-round communication. Read-only transactions in ChainReaction [28] can induce more than one-round communication as well as abort and retry, resulting in more communication. Eiger-PS [12] provides fast transactions and satisfies *process-ordered serializability* [12], stronger than causal consistency; yet in addition to the request-response of a transaction, each client periodically communicates with every server. Our Theorem 1 explains Eiger-PS's additional communication. COPS-SNOW [12] provides fast read-only transactions but writes can only be performed outside a transaction; moreover, any read-only transaction in COPS-SNOW is visible, complying with our Theorem 1 and Theorem 2. If each server stores a copy of all objects, then SwiftCloud [20] provides fast read-only transactions. However, it is not clear how such a storage can scale well with the growth of data given a single server. SwiftCloud considers the storage of a full copy among multiple servers and its resulting parallelism is an orthogonal issue [20].

5.3 Impossibility results

Existing impossibility results on storage systems have typically considered stronger consistency properties than causality or stronger progress conditions than eventual visibility. Brewer [10] conjectured the CAP theorem that no implementation guarantees *consistency*, and *availability* despite *partitions*. Gilbert and Lynch [11] formalized and proved Brewer's conjecture in *partially synchronous* systems. They formalized consistency by *atomic* objects [9] (which satisfy *linearizability*

[8], stronger than causal consistency). Considering a storage implemented by *data centers* (clusters of servers), if any value written is *immediately* visible to the reads at the same data center (as the write), and a client can access different objects at different data centers, Roohitavaf et al. [32] proved the impossibility of ensuring causal consistency and availability despite partitions. Their proof (as well as the proof of the CAP Theorem) rely on message losses. Lu et al. [12] proved the SNOW theorem, saying that fast *strict serializable* transactions [23, 24] (satisfying stronger consistency than causal consistency) are impossible. Their proof assume writes that are also fast, and does not imply our proof of impossibility results.

Mahajan et al. [33], Attiya et al. [34] as well as Xiang and Vaidya [35] proposed related notions of causal consistency based on the events at servers (rather than clients) motivated by replication schemes (an issue orthogonal to the problem considered in this paper). More specifically, Mahajan et al. [33] proved the CAC theorem that no implementation guarantees *one-way convergence*,[6] availability, and any consistency stronger than *real time causal consistency* assuming infinite local clock events and arbitrary message loss. Attiya et al. [34] proved that a *replicated* store implementing multi-valued registers cannot satisfy any consistency strictly stronger than *observable causal consistency*.[7] Xiang and Vaidya [35] proved that for *replica-centric causal consistency*, it is necessary to track down writes.

5.4 Transactional memory

In the context of transactional memory, if the implementation of a read-only operation (in a transaction) writes a base shared object, then the read-only operation is said to be *visible* and *invisible* otherwise [36]. Known impossibility results on invisible reads of TM assume stronger consistency than causal consistency. Attiya et al. [37] showed that no TM implementation ensures strict serializability, *disjoint-access parallelism* [37][8] and uses invisible reads, the proof of which shows that if writes are frequent, then a read can miss some write forever. Peluso et al. [38] considered any consistency that respects the real-time order of transactions (which causal consistency does not necessarily respect), and proved a similar impossibility result. Perelman et al. [39] proved an impossibility result for a multi-version TM implementation with invisible read-only transactions that ensures strict serializability and maintains only a necessary number of versions, the proof of which focuses on garbage collection of versions. None of the results or proofs above imply our impossibility results.

6 Concluding Remarks

Our impossibility results establish fundamental limitations on the performance on transactional storage systems. The first impossibility basically says that fast read-only transactions are impossible in a general setting where writes can also be performed within transactions. The second impossibility says that in a setting where all transactions are read-only, they can be fast, but they need to visible. A system like COPS-SNOW [12] implements such visible read-only transactions that leave traces when they execute, and these traces are propagated on the servers during writes. (For completeness, we sketch in Appendix D a variant algorithm where these traces are propagated asynchronously, i.e., outside writes).

[6] A progress condition based on the communication between servers.

[7] The definitions of observable causal consistency given by Mehdi et al. and Attiya et al. [34, 22] are different.

[8] Disjoint-access parallelism [37] requires two transactions accessing different application objects to also access different base objects.

Clearly, our impossibilities apply to causal consistency and hence to any stronger consistency criteria. They hold without assuming any message or node failures and hence hold for failure-prone systems. For presentation simplicity, we assumed that servers store disjoint sets of objects, but our impossibility results hold without this assumption (Appendix C). Some design choices could circumvent these impossibilities like imposing a full copy of all objects on each server (as in SwiftCloud [20]), periodic communication between servers and clients (as in Eiger-PS [12]), or transactions that abort and retry (as in ChainReaction [28]), [9] Each of these choices clearly hampers scalability.

We considered an asynchronous system where messages can be delayed arbitrarily and there is no global clock. One might also ask what happens with synchrony assumptions. If we assume a fully synchronous system where message delays are bounded and all processes can access a global accurate clock, then our impossibility results can be both circumvented. We give such a timestamp-based algorithm in Appendix D. If we consider however a system where communication delays are unbounded and all processes can access a global clock, then only our Theorem 1 holds. In this sense, message delay is key to the impossibility of fast read-only transactions, but not to the requirement that they need to be visible, in the restricted model where all transactions are read-only. In this restricted model, our timestamp-based algorithm of Appendix D can also circumvent Theorem 2 if we assume a global clock.

[9] Multiple versions (allowed to be returned in a transaction) do not circumvent our impossibility results as an infinite number of versions would be necessary.

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A Proof of Theorem 1

A.1 Definition of One-Version Property

In Section 2, we required that a server returns one and only one value which has been written to an object, a property we we call one-version, which we define below. The formal definition is necessary because (1) there are a lot of possibilities for message m to return value v, e.g., $m = v$, or $m = v$ XOR c, or $m = v + c$ for some constant c; and (2) if messages m_1 and m_2 are from two different servers P_X and P_Y and $m_1 = (x, \text{first 8 bits of } z \text{ XOR } c)$, $m_2 = (y, \text{other bits of } z \text{ XOR } c)$, where z is a value written to another object Z, then (m_1, m_2) can return more values x, y, z than expected. The first issue calls for defining messages in a general manner; the second situation should be excluded (as it is not implemented by any practical storage system to the best of our knowledge). The formal definition addresses both issues.

To define the value included in a message in general, we have to measure the information *revealed* by events and messages. We consider the maximum amount of information that any algorithm can output according to the given input: events and messages. We then restrict the class of algorithms any correct implementation may provide (to the client-side). For example, an algorithm that outputs 1 regardless of the input should be excluded. Hence two definitions, one on algorithms used to reveal information and one on information indeed revealed, are presented before the definition of one-version property.

Definition 8 (Successful algorithms)**.** Consider any algorithm, denoted by $\mathcal{A}$, whose input is some information i_E (events and messages) of execution E. The output of $\mathcal{A}$ is denoted by $\mathcal{A}(i_E)$. We say that $\mathcal{A}$ is *successful*

- If $v \in \mathcal{A}(i_{E^v})$, then in E^v, $w(a)v$ occurs; and
- For any value u, let E^u be the resulting execution where $w(a)v$ is replaced by $w(a)u$. Then $u \in \mathcal{A}(i_{E^u})$.

Definition 9 (Information revealed)**.** Consider execution E, client C and C's invocation I of some transaction. Denote by M any non-empty subset of message receiving events that occur at C (including message contents) during I. We say that M *reveals* $(n_2 - n_1)$ version(s) of an object a if

- Among all successful algorithms whose input is $v_{C,I}$, n_1 is the maximum number of values in the output that are also values written to a before the start of I;
- Among all successful algorithms whose input is $v_{C,I}$ and M, n_2 is the maximum number of values in the output that are also values written to a before the end of I;

where $v_{C,I}$ is C's *view*, or all events that have occurred at C (including the message content if an event is message receiving), before the start of I.

Definition 10 (One-version property)**.** Consider any execution E, any client C and C's invocation I of an arbitrary transaction T with non-empty read set R. (T is general here in that T may contain only a single read, i.e., the write set is empty and $|R| = 1$). For any non-empty set of servers A, let $\Lambda_{I,A} = R \cap \{\text{objects stored on } P | \forall P \in A\}$ and denote by $M_{I,A}$ the events of C receiving messages from any server in A (including message contents) during I. Then an implementation satisfies one-version property if

- $\forall E, \forall I, \forall A$, $M_{I,A}$ reveals at most one version for each object in $\Lambda_{I,A}$, and no version of any object not in $\Lambda_{I,A}$; and

- $\forall E, \forall I$, when A includes all servers, then $M_{I,A}$ reveals exactly one version for each object in R, and no version of any object not in R.

(If $M_{I,A}$ reveals exactly one version of an object a, we may also specify the version v and say that $M_{I,A}$ reveals v.)

A final remark is on the relation with the property of fast transactions. Naturally, when we consider the maximum amount of information revealed by a transaction, we have to consider all message receiving events at the client-side. As one-version property is defined in general here (independent from the property of fast transactions), there can be multiple message receiving events during a transaction. The formal definitions above consider the set of all these events rather than individual ones separately (i.e., what one message can reveal). This general definition is necessary to disallow implementations equivalent to fast transactions to bypass our results. Consider an equivalent implementation in a transaction of which a server splits its message to several ones and sends them to a client where each message reveals one version. Such implementation does not conform to our requirement on servers in Section 2 yet is however not excluded by a definition considering only individual messages, showing the necessity of a general definition as we present above.

A.2 Construction of E_{imp}

The construction of E_{imp} is based on the following notations and execution E_{prefix}. We denote by P_X the server which stores object X, and P_Y the server which stores object Y. Let E_{prefix} be any execution where X and Y have been written at least once and some values of X and Y have been visible. Denote by t_{start} when some values of X and Y have been visible in E_{prefix}. Then we construct execution E_{imp} starting from t_{start}. In E_{imp}, client C_w does transaction $WOT = (w(X)x, w(Y)y)$ which starts at some time $t_w > t_{start}$, while all other clients do no transaction. For E_{imp}, since t_w, WOT is the only transaction. However, for *any* positive number k, we show that k messages have to be sent and received after t_w and before x and y are visible. Since k can be any positive number, then k essentially goes to infinity.

More specifically, we show that no matter how many k messages have been sent and received, an additional message is necessary for x and y to be visible. I.e., our construction is by mathematical induction on the number k of messages, summarized in Proposition 1, Proposition 2 and Proposition 3. Each case k (Proposition 2) is a property of E_{imp} after k messages have been sent and received in the visibility of values written by WOT: if one read-only transaction ROT is added, then the values written by WOT cannot yet be returned to ROT. Here are some notations for ROT and the message schedule during ROT which we use in the statement of case k. Let C_r be the client which requests $ROT = (r(X)*, r(Y)*)$; C_r has requested no transaction before. By Definition 5, for any ROT, we schedule messages such that every message which C_r sends to either $P \in \{P_X, P_Y\}$ during ROT arrives at the same time t_P at P. After t_P and before P has sent one message to C_r (during ROT), P receives no message and any message sent by P to a process other than C_r is delayed to arrive after ROT ends. For either P, we denote these messages which P sends to C_r after t_P (during ROT) by $m_{resp,P}$. The message schedule of ROT such that P receives no message during ROT is also assumed in Lemma 2. Each case $k > 1$ is accompanied by a preliminary (Proposition 1) on the necessity of an additional kth message, while the base case is a special case for which two additional messages are necessary (Proposition 3). To deal with the special case, we index these messages starting from 0: $m_0, m_1, \ldots, m_{k-1}, m_k$ (but our base case is still the case where $k = 1$). As shown in Proposition 3, P_X and P_Y send m_X and m_Y after t_w that precede some message which arrive at P_Y and P_X respectively. We define m_0 and m_1 as follows so that the base case is the case where $k = 1$ defined in Proposition 1: one server between P_X and P_Y sends $m_0, m_0 \in \{m_X, m_Y\}$ before receiving any message which is preceded by m_1 for $\{m_0, m_1\} = \{m_X, m_Y\}$. We refer to Definition 6 for the formal definition on the relation of one message *preceding* another used in our Proposition 1, Proposition 2 and Proposition 3.

Proposition 1 (Additional message in case k)**.** *In E_{imp}, $m_0, m_1, \ldots, m_{k-1}$ have been sent. Let D_{k-1} be the source of m_{k-1}. Let $\{D_{k-1}, D_k\} = \{P_X, P_Y\}$. Let T_{k-1} be the time when the first message preceded by m_{k-1} arrives at D_k. After T_{k-1}, D_k must send at least one message m_k that precedes some message which arrives at D_{k-1}.*

Proposition 2 (Case k)**.** *In E_{imp}, $m_0, m_1, \ldots, m_{k-1}, m_k$ have been sent. Then for any t before T_k, if C_r starts ROT before t and $t_{D_{k-1}} = t$, then ROT may not return x or y.*

Proposition 3 (Additional message in the base case)**.** *After t_w, any $P \in \{P_X, P_Y\}$ must send at least one message that precedes some message which arrives at Q for $\{P, Q\} = \{P_X, P_Y\}$.*

A.3 Proof of Theorem 1

Before the proof of the base case and the inductive step from case k to case $k + 1$, we prove a helper lemma, Lemma 2. Lemma 2 is helpful for the proof of both the base case and case k, and thus proved additionally to avoid repetition. We also show a property of write-only transactions in Lemma 1. We refer to the main paper for its formal statement. As Lemma 2 is based on Lemma 1, we prove the latter first.

Proof of Lemma 1. By contradiction. Suppose that for some execution E_{imp} and some read-only transaction *ROT*, *ROT* returns (x^*, y) for some $x^* \neq x$, or (x, y^*) for some $y^* \neq y$. By symmetry, we need only to prove the former. As *ROT* returns (x^*, y), by causal consistency, for C_r, there is serialization $\mathcal{S}$ that orders all C_r's transactions and all transactions including a write such that the last preceding writes of X and Y before *ROT* in $\mathcal{S}$ are $w(X)x^*$ and $w(Y)y$ respectively. Therefore $\mathcal{S}$ must order *WOT* before $w(X)x^*$. However, if we extend E_{imp} with C_r requesting another read-only transaction ROT_2, then by progress, some ROT_2 must return (x, y). As ROT_2 occurs after *ROT*, $\mathcal{S}$ must order ROT_2 after *ROT* and then the last preceding writes of X and Y before ROT_2 in $\mathcal{S}$ cannot be $w(X)x$ and $w(Y)y$ respectively, contradictory to the property of causal consistency. □

Lemma 2 (Communication prevents latest values)**.** *Suppose that E_{imp} has been extended to some time A and there is no other write than contained in WOT since t_{start}. Let $\{P, Q\} = \{P_X, P_Y\}$ where P can be either P_X or P_Y. Given P, assume that for some time $B > A$ and any $t_P \in [A, B)$, if C_r starts ROT before t_P, then ROT may not return x or y. We have:*

1. *After B, P must send at least one message which precedes some message that arrives at Q;*
2. *Let t be the time when Q receives the first message which is preceded by some message which P sends after B. For any $\tau \in [A, t)$, if C_r starts ROT before τ and $t_Q = \tau$,[10] then ROT may not return x or y.*

Proof of Lemma 2. We prove the first statement by contradiction. Suppose that after B, P sends no message that precedes any message that arrives at Q. Let t_s be the latest time before B such that P sends a message that precedes some message which arrives at Q in E_{imp}. After t_s, we extend E_{imp} into two different executions E_1 and E_2. Execution E_2 is E_{imp} extended without any

[10] If needed, by the asynchronous communication, we may delay t after *ROT* ends to respect the message schedule of *ROT* that Q receives no message during *ROT*.

transaction. Thus x and y are visible after some time t_{ev}. Based on our assumption, $t_{ev} \geq B$. In E_2, C_r starts ROT after t_{ev}. In E_1, C_r starts ROT after t_s (and before B) and some $t_P \in [A, B)$. We delay any message which P sends after t_s in E_1. Furthermore, in both E_1 and E_2, $t_Q > t_{ev}$. According to our assumption, after t_s, P does not send any message which precedes some message that arrives at Q in E_2. As we delay the messages which P sends after t_s in E_1, thus before t_Q, Q is unable to distinguish between E_1 and E_2. After t_Q (inclusive), according to the message schedule of ROT, by the time when Q has sent one message to C_r during ROT, Q is still unable to distinguish between E_1 and E_2. In E_2, since C_r starts ROT after t_{ev}, ROT returns (x, y) by progress. The client-side algorithm $\mathcal{A}$ of C_r to output the return value of ROT is a successful algorithm. Since given $m_{resp,P}$ and $m_{resp,Q}$, $\mathcal{A}$ outputs (x, y), then by one-version property, $m_{resp,Q}$ reveals one and only one between x and y. (Otherwise, if $m_{resp,Q}$ can reveal another value v other than x and y, then we can obtain a successful algorithm which outputs x, y, v given $m_{resp,P}$ and $m_{resp,Q}$, violating one-version property.) By Q's indistinguishability between E_1 and E_2, in E_1, $m_{resp,Q}$ reveals one and only one between x and y. W.l.o.g., let $m_{resp,Q}$ reveal x. By the construction of E_{prefix}, the return value of ROT in E_1 cannot include $\perp$. As C_r has not requested any transaction before, then in E_1, the return value depends solely on $m_{resp,P}$ and $m_{resp,Q}$. As the client-side algorithm $\mathcal{A}$ is successful, thus $\mathcal{A}$ cannot output a value other than x for object X. As a result, ROT returns x in E_1. A contradiction to the assumption that if $t_P \in [A, B)$ (which matches E_1), then ROT may not return x or y.

We prove the second statement also by contradiction. Suppose that in some E_{imp}, for some $\tau \in [A, t)$, some ROT such that $t_Q = \tau$ returns x or y. By Lemma 1, ROT returns (x, y). Then we construct E_{old} which is the same as E_{imp} except that in E_{old}, ROT starts before B. In E_{old}, let $t_P \in (t_s, B)$ and let $t_Q = \tau$; all messages sent by P after t_s are delayed. Thus Q is unable to distinguish between E_{old} and E_{imp} by the time when Q has sent one message to C_r (for ROT). Since ROT returns (x, y) in E_{imp}, then $m_{resp,Q}$ reveals x or y in E_{old}. By the construction of E_{prefix}, the return value of ROT in E_{old} cannot include $\perp$. As C_r has not requested any transaction before, then in E_{old}, the return value depends solely on $m_{resp,P}$ and $m_{resp,Q}$, which must include x or y. A contradiction to the assumption in the statement of the lemma. □

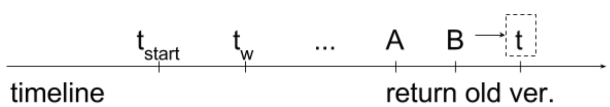


Figure 6: Timeline in Lemma 2

As illustrated in Figure 6, Lemma 2 is based on an assumption that before B, old versions are returned (if ROT is appropriately added), shows that B can be prolonged to time t. The proof of Lemma 2 relies on fast read-only transactions. What remains is the complete proof of Theorem 1, which proves Proposition 1, Proposition 2 and Proposition 3 as well.

Proof of Theorem 1. We first prove Proposition 2 for any positive k by mathematical induction and then show that E_{imp} indeed violates progress according to Definition 4.

By mathematical induction, we start with the base case, i.e., Proposition 3 and Proposition 2 for $k = 1$. Let $A = t_{start}$ and let $B = t_w$. By symmetry, we need only to prove Proposition 3 for $P = P_X$. We show that given P, for any $t_P \in [A, B)$, if C_r starts ROT before t_P, then ROT may not return x or y. For this ROT, as WOT has not yet started, $m_{resp,P}$ cannot reveal x or y. By one-version property, $m_{resp,P}$ reveals at most one version v_1 of X and $\{m_{resp,P}, m_{resp,Q}\}$ also reveals at most one version v_2 of X. Therefore $v_1 = v_2 \neq x$. As C_r has requested no transaction before, the return value of ROT solely depends on $m_{resp,P}$ and $m_{resp,Q}$. As the client-side algorithm of C_r for the return value of ROT is a successful algorithm, ROT returns $v_1 = v_2 \neq x$ for object X. (Due to E_{prefix}, ROT cannot return $\perp$.) Then by Lemma 1, ROT may not return x or y. Thus Lemma 2 applies. As a result, after $B = t_w$, P must send at least one message that precedes some message that arrives at Q. Therefore, Proposition 3 is true for either $P \in \{P_X, P_Y\}$. Following Proposition 3, recall the definition of m_0 and m_1. Let m_0 and m_1 be sent in E_{imp}. Recall that T_1 is the time when the first message preceded by m_1 arrives at D_0. According to Lemma 2, for any $t \in [A, T_1)$, if C_r starts ROT before t and $t_{D_0} = t$, then ROT may not return x or y, which proves Proposition 2 for $k = 1$.

We continue with the inductive step from case k to case $k+1$. For our assumption on k, let $A = T_{k-1}$, $B = T_k$, $P = D_{k-1}$ and $Q = D_k$. According to the definition of T_k, T_k is at least the time when m_k is received. By Proposition 1 for case k, m_k is sent at least after T_{k-1}. Therefore, $T_k > T_{k-1}$, or $B > A$. Thus Lemma 2 applies again. As a result, after T_k, $D_{k+1} = D_{k-1}$ must send at least one message m_{k+1} which precedes some message that arrives at D_k; let m_{k+1} be sent and then for any $t \in [T_{k-1}, T_{k+1})$, if C_r starts ROT such that $T_{D_k} = t$, then ROT may not return x or y, which proves Proposition 1 and Proposition 2 for case $k+1$. Therefore, we conclude Proposition 2 for any positive number k.

We show that E_{imp} violates progress by contradiction. Suppose that E_{imp} does not violate progress. As there is no other write since the start of WOT, then in E_{imp} there is finite time τ such that any read of object X (or Y) which starts at any time $t \geq \tau$ returns x (or y). We have shown that $T_{k+1} > T_k$ for any positive k. Thus for any finite time τ, there exists K such that for any $k \geq K$, $T_k > \tau$. Then there exists some k for which C_r starts ROT at some $t \geq \tau$ and $t_{D_{k-1}}$ is less than T_k. By Proposition 2, ROT may not return x or y. A contradiction. Therefore we find an execution E_{imp} where two values of the same write-only transaction can never be visible, violating progress. □

B Proof of Theorem 2

In this section, as a proof of Theorem 2, we show that if some implementation provides invisible read-only transactions, then we reach a contradiction. In other words, we show that for every implementation that provides fast read-only transactions, read-only transactions are *visible*.

B.1 Visible transactions

From Definition 7, a read-only transaction T is not invisible if for some client C and C's invocation I of T, some execution E (until I) can be continued arbitrarily and every execution E^- without I is different from E in addition to the message exchange with C (during the time period of I). We note that in this case, T does not necessarily leave a trace on the storage. It is possible that for some invocation I of T, some execution E (until I) can be continued arbitrarily and there is some execution E^- without I which is the same as E except for the message exchange with C (during the time period of I).

This motivates us to define the notion of being visible stronger than that of being not invisible, in Definition 11 below. In Definition 11, we assume that (1) for each object, some non-$\perp$ value has been visible; (2) the client C which invokes I has not done any operation before I; and (3) during I, C sends exactly one message m to the servers involved which receive m at the same time, while after the reception of m, all servers receive no message before I ends (but still respond to C). We note that even under these assumptions, Definition 11 still shows a strictly stronger notion than being not invisible.

Definition 11 (Visible transactions)**.** We say that transaction T is visible if for any invocation I of T, any execution E (until I) can be continued arbitrarily and every execution E^- without I is different from E in addition to the message exchange with the client which invokes I (during the time period of I).

Clearly, the definition of visible transactions does not yet quantify the difference between E and E^-, or show how much information is exposed by a visible transaction. In this proof, we quantify the exposed information by proving Proposition 4. Like Definition 11, we assume in Proposition 4 that (1) for each object, some non-$\perp$ value has been visible; (2) the clients $\mathcal{D}$ which invoke S_T have not done any operation before S_T; and (3) during S_T, every client C in $\mathcal{D}$ sends exactly one message m to the servers involved which receive m at the same time, while after the reception of m, all servers receive no message before S_T ends (but still respond to C). The property described in Proposition 4 is a strictly stronger variant of visible transactions. (To see this, one lets $I \in S_1$ and $I \notin S_2$.) Therefore, if we prove Proposition 4, then we also prove Theorem 2.

Proposition 4 (Stronger variant of visible transactions)**.** *Given any causally consistent storage system that provides fast read-only transactions, for some read-only transaction T, for any set $\mathcal{D}$ of clients and $\mathcal{D}$'s set S_T of concurrent invocations*[11] *of T, for any subset $S_1 \subseteq S_T$, any execution E_1 where only S_1 is invoked (the prefix until S_1) can be continued arbitrarily and every execution E_2 where only S_2 is invoked is different from E_1 in addition to the message exchange with $\mathcal{D}$ (during the time period of S_T) for any subset $S_2 \subseteq S_T$ where $S_2 \neq S_1$.*

To see that this variant quantifies the exposed information, we count the number of possibilities of these executions that are the same to all clients except for $\mathcal{D}$, with a subset of clients ss's invocations Inv_{ss} of T at the same time where $ss \subseteq \mathcal{D}$. If Proposition 4 is true for $S_T = Inv_{\mathcal{D}}$, then the number of possibilities is lower-bounded by the number *num* of subsets of $\mathcal{D}$, implying the amount of difference on the message exchange among these executions.

B.2 Executions

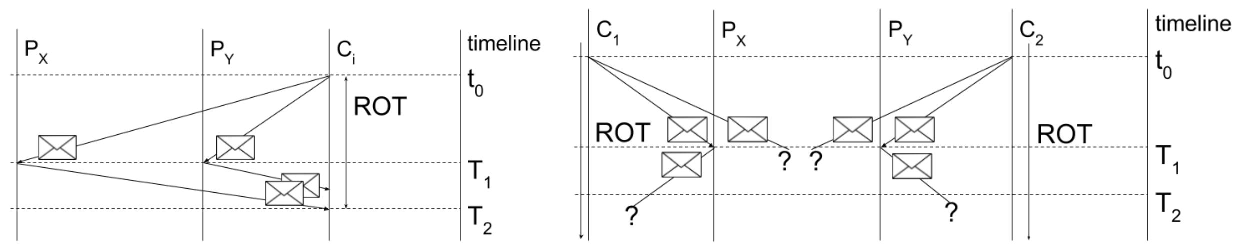
Our proof is by contradiction. Suppose that for any read-only transaction T, for some set $\mathcal{D}$ of clients and $\mathcal{D}$'s set S_T of concurrent invocations of T, for some subset $S_1 \subseteq S_T$, some execution E_1 where only S_1 is invoked (the prefix until S_1) can be continued arbitrarily but still there exists some execution E_2 where only S_2 is invoked and which is the same as E_1 except for the message exchange with $\mathcal{D}$ during the time period of S_T for some subset $S_2 \subseteq S_T$ where $S_2 \neq S_1$.

We thus construct two executions E_1 and E_2 following our idea of quantifying information previously. We first recall our construction of the set S_T. (We are allowed to do so, as the set is assumed so in the assumption for Proposition 4.) Let S_T be the invocations of $ROT = (r(X)*, r(Y)*)$ each of which is invoked by one client in $\mathcal{D}$ at the same time t_0. Furthermore, we consider S_T that are performed as follows. By fast read-only transactions, all messages which a client in $\mathcal{D}$ sends during ROT arrive at P_X and P_Y respectively at the same time. Let T_1 denote this time instant and let T_2 be the time when ROT eventually ends. During $[T_1, T_2]$, by fast read-only transactions, P_X and P_Y receive no message. If there is any such message, they are delayed to at least after T_2 but eventually arrive before τ_y.

Now in some E_1, only S_1 is invoked but every other detail about the execution of S_T above remains the same. We continue E_1 with client C performing two writes $w(X)x$ and $w(Y)y$ after T_2 to establish $w(X)x \rightsquigarrow w(Y)y$ according to Definition 1. Moreover, after T_2, the clients in $\mathcal{D}$

[11] Some invocations are said to be concurrent here if the time period between the start and end of these invocations are the same (stronger than the common definition of concurrency).

do not invoke any operation. (We are allowed to do so, as E_1 can be continued arbitrarily in our assumption for contradiction.) According to our assumption for contradiction, some E_2 is the same as E_1 except for the message exchange with $\mathcal{D}$ during the time period of S_T, although in E_2, only S_2 is invoked and $S_2 \neq S_1$. Both executions are illustrated in Figure 7a before the two writes. In both executions, y is eventually visible. We denote by τ the time instant after which y is visible in both executions.



(a) Message schedule of E_i

(b) Message schedule of $E_{1,2}$

Figure 7: Construction and extension of E_i

For $i \in \{1, 2\}$, let $\mathcal{D}_i$ be the subset of clients which invoke S_i in E_i. Let C_i be any client in $\mathcal{D}_i$. As E_2 is the same as E_1 except for the message exchange with $\mathcal{D}$ during the time period of S_T, w.l.o.g., we assume that $\mathcal{D}_1 \backslash \mathcal{D}_2 \neq \emptyset$. We denote by C_r any client in $\mathcal{D}_1 \backslash \mathcal{D}_2$ hereafter.

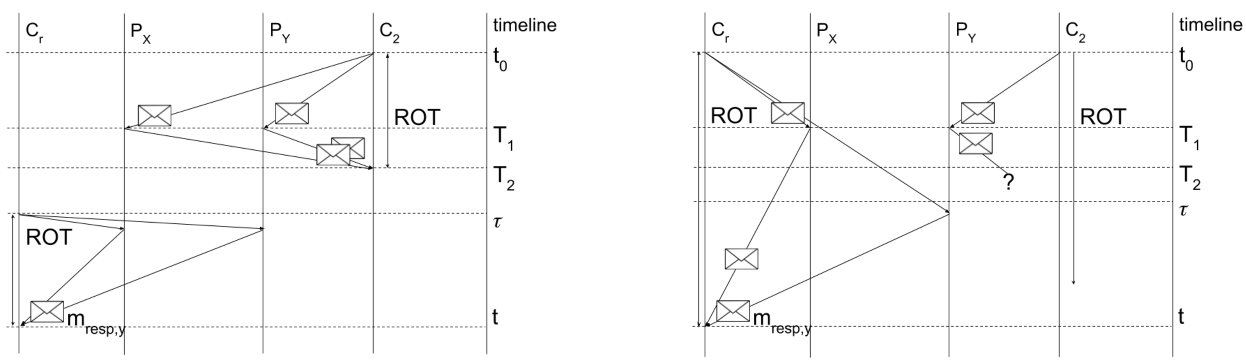
We next construct an execution $E_{1,2}$ based on E_1 and E_2 to help our proof. Our goal is to let $E_{1,2} = E_1 = E_2$ except for the communication with $\mathcal{D}$ (during the time period of S_T) until the same time τ. In $E_{1,2}$, every client in $\mathcal{D}_1 \cup \mathcal{D}_2$ invokes ROT at t_0. As illustrated in Figure 7b, while every client $C_1 \in \mathcal{D}_1$ invokes ROT, P_X receives the same message from C_1 at the same time T_1 and no other message during $[T_1, T_2]$, and sends the same message to C_1 at the same time as in E_1; C_1 sends the same message to P_Y at the same time as in E_1, the reception of which may however be delayed by a finite but unbounded amount of time (see below). Similarly, while every client $C_2 \in \mathcal{D}_2$ invokes ROT, P_Y receives the same message from C_2 at the same time T_1 and no other message during $[T_1, T_2]$, and sends the same message to C_2 at the same time as in E_2; C_2 sends the same message to P_X at the same time as in E_2, the reception of which may however be delayed by a finite but unbounded amount of time (see below). For those clients in $\mathcal{D}_1 \cap \mathcal{D}_2$, the messages which are said to be possibly delayed still arrive at T_1 and follow both the message schedules of $\mathcal{D}_1$ and $\mathcal{D}_\in$ above. For the other clients, the messages are indeed delayed by a finite but unbounded amount of time. Furthermore, any message which P_X or P_Y sends to a process in $\mathcal{D}$ during $[T_1, T_2]$ is delayed to arrive at least after T_2. Thus by T_2, P_X is unable to distinguish between E_1 and $E_{1,2}$ while P_Y is unable to distinguish between E_2 and $E_{1,2}$. As a result, the first message $m_{X,1}$ which P_X sends after T_2 in $E_{1,2}$ is the same message as in E_1, and the first message $m_{Y,1}$ which P_Y sends after T_2 in $E_{1,2}$ is the same message as in E_2.

According to our assumption for contradiction, E_1 and E_2 are the same except for the communication with $\mathcal{D}$ during $[t_0, T_2]$. In other words, E_1 and E_2 are the same regarding the message exchange among servers and message exchange between any server and any client after T_2. Therefore, the first message which P_X sends after T_2 in E_2 is also $m_{X,1}$ and the first message which P_Y sends after T_2 in E_1 is also $m_{Y,1}$. Therefore, the message exchange among servers in $E_{1,2}$ continues in the same way as in E_1 as well as E_2 after T_2. Since $\mathcal{D}$ does not invoke any operation after T_2 in both executions, then after T_2, no client can distinguish between E_1 and E_2 and therefore the message exchange between any server and any client in $E_{1,2}$ continues also in the same way as in E_1 as well as E_2 after T_2. Then even if the delayed messages in $E_{1,2}$ do not arrive before τ,

$E_{1,2} = E_1 = E_2$ except for the communication with $\mathcal{D}$ (during the time period of S_T) until the same time τ. We reach our goal as stated previously.

B.3 Proof

Our proof starts with the extension of E_2 and $E_{1,2}$ since the time instant τ. We show that in a certain extension, P_Y is unable to distinguish between E_2 and $E_{1,2}$ and thus returns a new value, which breaks causal consistency. As we reach a contradiction here, we show the correctness of Proposition 4 as well as that of Theorem 2. We also have two remarks on the proof of Proposition 4. First, the proof relies on the indistinguishability of servers between executions, implying that fast read-only transactions have to "write" to some server to break the indistinguishability (i.e., "writing" to a client without the client forwarding the information to a server is not an option). Second, recall that to quantify the exposed information of read-only transactions, we count the number of possibilities of these executions that are the same to all clients except for $\mathcal{D}$, with a subset ss of $\mathcal{D}$'s invocations Inv_{ss} of T at the same time. Now that the proof shows that Proposition 4 is indeed true for $S_T = Inv_{\mathcal{D}}$, then the number of possibilities is lower-bounded by the number 2^n where $n = |\mathcal{D}|$, implying that each fast read-only transaction in S_T contributes at least one bit in the message exchange.[12]



(a) Extension of E_2

(b) Extension of $E_{1,2}$.

Figure 8: Extension of two executions

Proposition 5 (Contradictory execution)**.** *Execution $E_{1,2}$ can violate causal consistency.*

Proof of Proposition 5. We first extend E_2 and $E_{1,2}$ after τ. As illustrated in Figure 8, we let any client C_r in $\mathcal{D}_1 \backslash \mathcal{D}_2$ start *ROT* immediately after τ in E_2. Then in both E_2 and $E_{1,2}$, we schedule the message sent from C_r to P_Y during C_r's *ROT* to arrive at the same time after τ, and by fast read-only transactions and asynchrony, $\exists t$, during $[\tau, t]$, P_Y receives no message but still responds to C_r. Based on our extension of E_2 and $E_{1,2}$, by t, P_Y is unable to distinguish between E_2 and $E_{1,2}$. We also schedule any message which P_X or P_Y sends to C_r arrives at the same time t. Denote the message which C_r receives from P_Y at t by $m_{resp,Y}$, which is thus the same in E_2 and $E_{1,2}$. Denote by $m_{resp,X}$, the message which C_r receives from P_X at t, which can be different in E_2 and $E_{1,2}$.

[12] The contribution is computed according to the information theory and coding theory. Consider $\mathcal{X}$ as a random variable that takes values in all these 2^n executions. Assume that $\mathcal{X}$ takes any one with equal probability. Then the entropy of $\mathcal{X}$ is n bits. According to the coding theory, depending on how the messages exchanged in these executions code S_T, each fast read-only transaction may use more than one bits.

We now compute the return value of ROT in E_2 and $E_{1,2}$. By progress, in E_2, C_r's ROT returns y for Y. By one-version property, $m_{resp,Y}$ reveals exactly one version of Y, and $m_{resp,X}$ reveals no version of Y. Since $m_{resp,X}$ reveals no version of Y, $m_{resp,Y}$ cannot reveal a version of Y different from y. In other words, $m_{resp,Y}$ must reveal y. In $E_{1,2}$, $m_{resp,X}$ cannot reveal x as $w(X)x$ starts after T_2. Then $m_{resp,X}$ must reveal some value $x^* \neq x$. As $m_{resp,Y}$ has already revealed y, messages $\{m_{resp,X}, m_{resp,Y}\}$ cannot reveal other versions of X or Y. As in $E_{1,2}$, some values of X and Y have been visible, then the return value cannot be $\perp$; thus the return value of ROT in $E_{1,2}$ is (x^*, y).

We show that the return value (x^*, y) in $E_{1,2}$ violates causal consistency by contradiction. Suppose that $E_{1,2}$ satisfies causal consistency. Then by Definition 3, we can totally order all C_r's operations and all write operations such that the last preceding writes of X and Y before C_r's ROT are $w(X)x^*$ and $w(Y)y$ respectively. Since $x \rightsquigarrow y$, then $w(X)x$ must be ordered before $w(Y)y$. This leads $w(X)x^*$ to be ordered after $w(X)x$. We now extend $E_{1,2}$ so that C_r invokes $ROT_1 = (r(X)*, r(Y)*)$ after x is visible, which returns value (x, y) by Definition 4. According to Definition 3, the last preceding write of X before ROT_1 must be $w(X)x$. However, $w(X)x^*$ has already been ordered after $w(X)x$ and thus the last preceding write of X before ROT_1 is $w(X)x^*$. A contradiction. □

C Storage Assumptions

For presentation simplicity, we made an assumption that servers store disjoint sets of objects. In this section, we show how our results apply to the non-disjoint case. A general model of servers' storing objects can be defined as follows. Each server still stores a set of objects, but no server stores all objects. For any server S, there exists object o such that S does not store o. In this general model, when a client reads or writes some object o, the client can possibly request multiple servers all of which store o. W.l.o.g., we assume that when client C accesses o, C requests all servers that store o.

C.1 Impossibility of fast transactions

We sketch here why Theorem 1 still holds in the general model. To prove Theorem 1, we still construct a contradictory execution E_{imp} which, to satisfy causality, contains an infinite number of messages and then violates progress. In E_{imp}, C_w issues a write-only transaction WOT which starts at time t_w and writes to all objects. In other words via WOT, C_w writes all objects. Since t_w, WOT is the only executing transaction.

The proof is still by induction on the number of messages. (Here k does not denote the number of messages; rather k denotes the number of asynchronous rounds of messages as shown by our inductive step.) We sketch the base case and the inductive step below as the main idea of the proof is the same as in Appendix A. We add an imaginary read-only transaction ROT which reads all objects to E_{imp}. Let n_{obj} be the number of objects read by ROT. For each server S, let $m_{resp,S}$ be the message of response from S during ROT if ROT is invoked after all values written by WOT are eventually visible. Let ss be the smallest set of servers such that $\{m_{resp,S} | \forall S \in ss\}$ reveals exactly n_{obj} versions. By one-version property, ss contains at least two servers. In the base case, we show that after t_w, there are at least two servers each of which sends some message that precedes some message which arrives at another server in ss. By contradiction. Suppose that at most one server sends such message. Since at most one server sends some message that precedes some message which arrives at another server in ss, then we assume that one server $R \in ss$ does not send any such message. We now make our ROT concrete. We let the request of ROT arrive earlier than t_w

at R and delay the request ROT at $ss\backslash\{R\}$. Based on our assumption on R, $ss\backslash\{R\}$ is unable to distinguish between the case where ROT starts before t_w and the case where ROT has not started at all. Then all values written by WOT are eventually visible. After that, we let the request of WOT arrive at $ss\backslash\{R\}$. By fast read-only transactions and our assumption for contradiction, after t_w, there can be no communication between R and $ss\backslash\{R\}$. Therefore $ss\backslash\{R\}$ returns values written by WOT; however R has to return some value not written by WOT. We note that $ss\backslash\{R\} \neq \emptyset$ and thus by Lemma 1, the return value of ROT must break causal consistency. A contradiction.

Therefore we conclude that after t_w, there are at least two servers each of which sends some message that precedes some message which arrives at another server in ss. Let S_{base} be the set of servers (whether in ss or not) which do so. Let M_{base} be the set of first messages which (1) a server in S_{base} sends and (2) precedes some message that arrives at another server in ss.

We now sketch the inductive step. Let P be the first server which receives some message m_P in M_{base}. Then from case $k = 1$ to case $k = 2$, we show that after the reception of m_P, at least one server sends some message that precedes some message which arrives at another server in ss. By contradiction. Suppose that after the reception of m_P, no server sends any such message. We now make our ROT concrete. We let the request of ROT arrive at one server R in ss before the reception of m_P, and let the request of ROT arrive at $ss\backslash\{R\}$ after all values written by ROT are visible. Considering the possibility that the request of ROT can arrive at another server before t_w, R returns at most one old version for each object R stores. However, by our assumption for contradiction, $ss\backslash\{R\}$ is unable to distinguish this case from the case where ROT starts after all values written by WOT are visible and therefore returns values written by WOT. We note that $ss\backslash\{R\} \neq \emptyset$ and thus by Lemma 1, the return value of ROT must break causal consistency. A contradiction.

Therefore we conclude that after the reception of m_P, at least one server sends some message that precedes some message which arrives at another server in ss. Let S_2 be the set of servers (whether in ss or not) which sends some message that precedes some message which arrives at another server in ss after receiving a message in M_{base}. Let M_2 be the set of first messages which (1) a server in S_2 sends and (2) precedes some message that arrives at another server in ss.

With an abuse of notations, let P be the first server which receives some message m_P in M_2. For case $k = 3$, we can similarly show that after the reception of m_P, at least one server sends some message that precedes some message which arrives at another server in ss. In this way, we add at least one message in each inductive step and also make progress in time which at the end goes to infinity. This completes our construction of E_{imp} as well as the proof sketch of Theorem 1 in the general model of servers' storing objects.

C.2 Impossibility of fast invisible transactions

We sketch here why Theorem 2 still holds in the general model. To prove Theorem 2, we consider any execution E_1 where some client C_r (which has not requested any operation before) starts transaction ROT which reads all objects at the time t_0 and before t_0, some values of X and Y have been visible. We continue E_1 with some client C executing writes which establishes a chain of causal relations. Let $O = \{o_1, o_2, \ldots, o_{n_{obj}}\}$ be the set of all objects. C executes writes $Wr = \{w(o)v | \forall o \in O\}$ so that $\forall k \in \mathbb{Z}$, $2 \leq k \leq n_{obj}$, $w(o_{k-1})v_{k-1} \rightsquigarrow w(o_k)v_k$.

Our proof is by contradiction. Suppose that transaction ROT is invisible. Then no matter how E_1 is scheduled, there exists some execution E_2 such that E_2 is the same as E_1 except that (1) C_r does not invoke ROT, and (2) the message exchange with C_r during the time period of ROT is different. Below we first schedule E_1 and then construct $E_{1,2}$. We later show $E_{1,2}$ violates causal consistency.

By fast read-only transactions, we can schedule messages such that the message which C_1 sends during ROT arrives at every server respectively at the same time. Let T_1 denote this time instant and let T_2 be the time when ROT eventually ends. During $[T_1, T_2]$, every server receives no message but still respond to C_r. Clearly all writes of C occur after T_2, while C_r does no operation after T_2. All delayed messages eventually arrive before all values written by C can be visible. In E_1, we denote the time instant after which all values written are visible by τ. Next we construct execution $E_{1,2}$ that is indistinguishable from E_1 to P_X and from E_2 to P_Y. The start of $E_{1,2}$ is the same as E_1 (as well as E_2) until t_0. At t_0, C_r still invokes ROT.

Before we continue the construction of $E_{1,2}$, we consider an imaginary ROT in E_2 which starts after τ. For each server S, let $m_{resp,S}$ be the message of response from S during this imaginary ROT if ROT is invoked after τ. Let ss be the smallest set of servers such that $\{m_{resp,S} | \forall S \in ss\}$ reveals exactly n_{obj} versions. By one-version property, there are at least two servers in ss. Let R be one server in ss such that $m_{resp,R}$ does not reveal v_1.

We now go back to our construction of $E_{1,2}$. We let all servers except for R receive the same message from C_r and send the same message to C_r at the same time as in E_1; C_r sends the same message to R at the same time as in E_1, the reception of which is however delayed by a finite but unbounded amount of time. In addition, during $[T_1, T_2]$, all servers receive no message as in E_1 (as well as E_2). Thus by T_2, all servers except for R are unable to distinguish between E_1 and $E_{1,2}$ while R is unable to distinguish between E_2 and $E_{1,2}$. Since $E_1 = E_2$ except for the communication with C_r during $[t_0, T_2]$, then $E_{1,2} = E_1 = E_2$ except for the communication with C_r during $[t_0, T_2]$.

Now based on the executions constructed above, we can similarly extend E_2 and $E_{1,2}$ after τ. We let C_r start ROT immediately after τ in E_2. In both E_2 and $E_{1,2}$, by fast read-only transactions, we schedule the message sent from C_r to R during C_r's ROT to arrive at the same time after τ, and $\exists t$, during $[\tau, t]$, R receives no message but still responds to C_r. Thus by t, R is unable to distinguish between E_2 and $E_{1,2}$.

We next compute the return value of ROT in $E_{1,2}$. By progress and the fact that $R \in ss$, in E_2, R returns some new values written by C, and then by indistinguishability, in $E_{1,2}$, R returns the same. However, in $E_{1,2}$, all servers except for R can only return some values which are written before Wr. W.l.o.g., the return value of ROT in $E_{1,2}$ includes some value $v_1^* \neq v_1$ for object o_1 and v_k for some object o_k. According to our assumption, $E_{1,2}$ satisfies causal consistency. By Definition 3, we can totally order all C_r's operations and all write operations in $E_{1,2}$ such that the last preceding writes of o_1 and o_k before C_r's ROT are $w(o_1)v_1^*$ and $w(o_k)v_k$ respectively. This leads $w(o_1)v_1^*$ to be ordered after $w(o_1)v_1$. However, if we extend $E_{1,2}$ so that C_r invokes $ROT_1 = (r(o_1)*, r(o_k)*)$ after τ, then ROT_1 returns value (v_1, v_k) and if we do total ordering of $E_{1,2}$ again, then the last preceding write of o_1 before ROT_1 must be $w(o_1)v_1$, which leads $w(o_1)v_1$ to be ordered after $w(o_1)v_1^*$, a contradiction. Therefore, we can conclude that $E_{1,2}$ violates causal consistency, which completes our proof sketch of Theorem 2 in the general model.

D Alternative Protocols

For completeness of our discussion in Section 6, we here sketch two implementations, one using asynchronous propagation of information among servers and one assuming the existence of a global accurate block.

D.1 Visible fast read-only transactions

We sketch below an algorithm $\mathcal{A}$ for fast read-only transactions. To comply with our Theorem 1, we restrict all transactions to be read-only. The goal of $\mathcal{A}$ is to better understand our Theorem

2. Theorem 2 shows that fast read-only transactions are visible. The intuition of Theorem 2 is that after a fast read-only transaction T, servers may need to communicate the information of T among themselves. However, it is not clear when such communication occurs. The COPS-SNOW [12] algorithm shows that the communication can take place during clients' requests of writes. Our algorithm $\mathcal{A}$ below shows that the communication can actually take place asynchronously. In addition, while COPS-SNOW guarantees a value to be visible immediately after its write, $\mathcal{A}$ guarantees only eventual visibility; thus a trade-off between the freshness of values and latency perceived by clients is also implied.

We sketch below first the data structure which each process maintains. All processes maintain locally their logical timestamps and update their timestamps whenever they find their local ones lag behind. They also move their logical timestamps forward when some communication with other processes is made. Every client additionally maintains the causal dependencies of the current operation (i.e., the operations each of which causally precedes the current one), called context. The maintenance of context is done in a similar way as COPS [16] and COPS-SNOW [12]. Every server needs to store the causal dependencies passed as an argument of some client' write. Every server additionally maintains a data structure called $OldTx$ for each object stored. We next sketch how writes and read-only transactions are handled.

- Every client sends its logical timestamp as well as context when requesting a write. A server stores the value written along with the server's updated logical timestamp, causal dependencies, and returns to the client.
- Every client C sends its logical timestamp as well as context when requesting a read-only transaction tx. A server first searches tx in $OldTx$, and returns a pre-computed value according to entry tx in $OldTx$ if $tx \in OldTx$. Otherwise, a server returns some value already observed by C (in its context) or some value marked as "visible".

We finally sketch how $OldTx$ is maintained and communicated (where asynchronous propagation mentioned in Section 6 takes place).

- After a server S responds to a client's write request of value w, S sends a request to every server which stores some value v such that $v \rightsquigarrow w$. Any server responds such request with its local $OldTx$ when v is marked as "visible".
- After S receives a response from all servers which store some value that causally precedes w, S stores their $OldTx$s into S's local one, chooses a value already observed by the client of tx or a value w^* which is written before w[13] for each transaction tx in $OldTx$ and marks w as "visible".

Any read-only transaction is stored and marked as "current" during its operation at any server. A "current" transaction T is put in $OldTx$ when some value w is "visible" and T has returned a value written before w of the same object.

Proof sketch of Correctness. Our algorithm $\mathcal{A}$ above provides fast read-only transactions. As every message eventually arrives at its destination (and therefore asynchronous propagation eventually ends), $\mathcal{A}$ satisfies progress. We can show that $\mathcal{A}$ satisfies causal consistency by contradiction.

[13] In order to choose a value correctly, in the algorithm, S actually sends a request after all values written before w are marked as "visible". Also, S does not choose a value for some tx which S has chosen before (which can happen when some value written before w is marked as "visible"). In this way, S can choose w^* as the last value written before w.

Suppose that some execution E violates causal consistency. Then E includes at least one read-only transaction. Assume that in E, for some client C, the ordering of all writes and C's transactions breaks causal consistency. Clearly, without any read-only transaction, we can order all writes in a way that respects causality. In addition, we can also order all writes of the same object according to the increasing timestamps of these writes and still respect causality. (We call the ordering of writes of the same object according to the timestamps by object relation. In addition, we say that two writes $w_1 \rightarrow w_2$, if (1) w_1 is before w_2 by object relation or by causal relation, or (2) $\exists$ some write w_3 such that $w_1 \rightarrow w_3$ and $w_3 \rightarrow w_2$.) Let to be any such ordering. We then add C's read-only transactions on to one by one. Let T be the first read-only transaction such that some to_1 exists which can include C's read-only transactions before T but for any to, T as well as C's read-only transactions before T cannot be placed in to property (i.e., to satisfy causal serialization.)

Let A be the set of such ordering to that can include C's read-only transactions before T and let to_1 be any ordering in A. We first claim a property of to_1: for any two reads $r(a)u, r(o)v^* \in T$, if in to_1, $\exists w(a)u^*$ such that $w(a)u$ is before $w(a)u^*$ and $w(a)u^*$ is before $w(o)v^*$, then $w(a)u^* \rightarrow w(o)v^*$ does not hold. For space limitation, we omit the proof of this claim. Second, based on the claim and to_1, we construct to_2 as follows. For any $r(a)u \in T$, consider $w(a)u^*$ as the first write of a such that (1) $w(a)u^*$ is after $w(a)u$, (2) some $w(o)v^*$ is after $w(a)u^*$ and (3) $r(o)v^* \in T$. We let W_u be the set of such write $w(o)v^*$ that is after $w(a)u^*$ and (3) $r(o)v^* \in T$. We then augment W_u by adding the precedence of each element according to relation $\rightarrow$, and we do this until no more write after $w(a)u^*$ in to_1 can be added. Let ss be the subsequence of to_1 which contains all writes in W_u. We move ss immediately before $w(a)u^*$. Below we verify that the resulting ordering to_m (not yet our goal to_2) falls in A. By the construction based on relation $\rightarrow$, to_m still respects causality and orders all writes of the same object according to the timestamps of these writes. We also verify that C's read-only transactions before T can be placed in to_m by contradiction: suppose that some read-only transaction T_0 before T finds the last preceding write of T_0 incorrect. As a result, T_0 must be after $w(a)u^*$ back in to_1; then $r(a)u^* \in T_0$; however, as T returns a value at least observed by C's previous operations, T cannot return u when T_0 has returned u^*, which gives a contradiction. Now that the move of ss creates no new pair $w(a)u$ and $w(o)v^*$ such that $r(a)u, r(o)v^* \in T$ and $w(o)v^*$ is after $w(a)u^*$ and $w(a)u^*$ is after $w(a)u$, then after a finite number of moves, we can construct an ordering $to_2 \in A$ such that for any $r(a)u \in T$, $W_u = \emptyset$. Finally, if we place T after the last write that corresponds some read in T in to_2, then we find all preceding writes of T are correct, a contradiction of our assumption. As a result, we must conclude that $\mathcal{A}$ satisfies causal consistency.

D.2 Timestamp-based implementation

The algorithm here relies on the assumption that all processes can access a global accurate clock and accurate timestamps:

- Before any client starts a transaction, the client accesses the clock and stamps the transaction with the current time;
- Every client sends the accurate timestamp while requesting a transaction;
- If an operation writes a value to an object, then the server that stores the object attaches the timestamp to the value;
- If an operation reads a value of an object, then the server that stores the object returns the value with the highest timestamp which is still smaller than the timestamp stamped by the client of the transaction.

Each transaction induces one communication round and is invisible. The algorithm guarantees progress as the clock makes progress.

If updates are only allowed outside transactions, then the algorithm satisfies causal consistency trivially as the accurate timestamp serializes all these individual writes, The algorithm thus circumvents our Theorem 2 no matter whether communication delays are bounded or not.

If general transactions are allowed, then the algorithm can be adapted to still satisfy causal consistency when the message delay is upper-bounded by time u. More specifically, a client imposes that every transaction is executed for time $2u$ and instead of comparing with the timestamp ts stamped by the client C, the server compares the timestamp of a value with $ts - 2u$ when responding to a read. All writes are still serialized, and these writes linked within the same transaction can be serialized at the same time. The algorithm thus circumvents our Theorem 1 when communication delays are bounded but a global accurate clock is accessible.