DTranx: A SEDA-based Distributed and Transactional Key Value Store with Persistent Memory Log

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ABSTRACT

Current distributed key value stores achieve scalability by trading off consistency. As persistent memory technologies evolve tremendously, it is not necessary to sacrifice consistency for performance. This paper proposes DTranx, a distributed key value store based on a persistent memory aware log. DTranx integrates a state transition based garbage collection mechanism in the log design to effectively and efficiently reclaim old logs. In addition, DTranx adopts the SEDA architecture to exploit higher concurrency in multi-core environments and employs the optimal core binding strategy to minimize context switch overhead. Moreover, we customize a hybrid commit protocol that combines optimistic concurrency control and two-phase commit to reduce critical section of distributed locking and introduce a locking mechanism to avoid deadlocks and livelocks.

In our evaluations, DTranx reaches 514.11k transactions per second with 36 servers and 95% read workloads. The persistent memory aware log is 30 times faster than the SSD based system. And, our state transition based garbage collection mechanism is efficient and effective. It does not affect normal transactions and log space usage is steadily low.

1. INTRODUCTION

There are numerous storage management systems, such as distributed RDBMS, NoSQL database, distributed file systems, and transactional key value stores. These systems offer different levels of transactional and data schema support. Distributed RDBMS provides strict data schema and full ACID properties with the price of low availability and efficiency. NoSQL databases and distributed file systems, such as Cassandra [16] and GFS [13], are scalable and highly available, but they often lack consistency support. Transactional key value stores, such as BigTable [7], sacrifice data schema flexibility, but they offer higher availability, superior performance, and better scalability.

In this paper we present DTranx, a SEDA-based distributed transactional key value store with persistent memory log. DTranx follows the SEDA [24] architecture to exploit the high concurrency in multi-core environments. SEDA organizes the software in a network of stages where stages contain both the application logic and communication channels. DTranx adopts lock free queues as the communication channels to reduce contention among threads. In addition, DTranx binds threads to physical cores to minimize the context switch overhead.

Unlike most existing key value stores, DTranx is fully ACID compliant supporting serializability. To serialize concurrent transactions, we adopt a hybrid of Optimistic Concurrency Control(OCC) and Two-Phase Commit(2PC) to narrow down the critical section of distribute locking to the commit time and enables parallel validation for high scalability. Furthermore, we avoid deadlocks and livelocks with a customized locking mechanism where transactions are aborted if shared lock requests are rejected and the exclusive lock requests are blocked for a timeout if not granted immediately. However, if the data is exclusively locked when the new exclusive lock requests come, the new requests are rejected immediately.

Moreover, DTranx integrates a modular Write-Ahead Log (WAL) which can be configured to use conventional SATA SSDs or Non-Volatile Memory(NVM) [23] technologies. Applying NVM in the WAL considerably cuts down the durability cost that most ACID-compliant systems suffer. A state transition mechanism to garbage collect(GC) WALs is also developed to reclaim the logs of the completed transactions. The garbage collection process does not affect normal transactions since old logs and the current appending log are not in the same file.

In summary, our contributions are as follows:

- Adopting SEDA concurrent architecture and employing the optimal core binding strategy;
- Customizing a hybrid commit protocol combining optimistic concurrency control and two-phase commit and introducing a locking mechanism to avoid deadlocks and livelocks;
- Adopting NVM using the Linux pmem library in the WAL of the distributed transactional system to reduce the persistence overhead and to offer durability;
- And, designing a state transition based garbage collection mechanism to efficiently reclaim increasing log space without affecting normal transactions.
2. BACKGROUND

Staged Event-Driven Architecture SEDA is a highly concurrent architecture, consisting of a network of event-driven stages connected by queues. A stage is an independent software module that manages a shared resource. For example, the lock service in transactional systems is a stage that maintains the locking information and handles lock requests. As shown in Figure 1, a stage is composed of an incoming event queue, a handler, and a thread pool. Besides the three core elements, SEDA adds a controller to adjust the thread pool size dynamically. The event handler sends events to another stage by invoking the enqueue operation on the incoming event queue of that stage. SEDA brings four benefits. First, it offers modularity and independent load management. Second, it facilitates debugging and performance analysis, which has always been a tough task for multi-threaded programs. Third, it optimizes the overall system performance by dynamically adjusting resource allocations, such as thread numbers among stages. Fourth, it enables batch request processing. For example, a database stage could write multiple keys at a time. However, SEDA requires nonblocking design of the event handler.

Optimistic Concurrency Control Concurrent control is the coordination of multiple concurrent accesses to the database and Philipp et al. [14] decomposed it into two major subproblems: read-write synchronization and write-write synchronization. There are pessimistic and optimistic approaches towards both subproblems. The pessimistic approach assumes the probability of access conflicts to be high and decides whether to restart at the start of transactions, such as two-phase locking. The optimistic approach assumes the probability of access conflicts to be low and decides whether to restart at the end of transactions. Specifically, Optimistic Concurrency Control (OCC) [14] consists of three phases: read, validation, and write. During the read phase, transactions read databases and store updated data in the buffer. Then, databases check whether the current transaction is in conflict with any concurrent operations. Finally, if it passes the validation phase, the current transaction proceeds to update the database states.

Two-Phase Commit Two-Phase Commit (2PC) is a classic commit protocol in the distributed environment that guarantees agreement among servers on the commit results. Moreover, once the agreement is reached, the commit results hold however the servers fail. There are two roles for the servers: coordinator and participant. During the first phase, coordinators initiate 2PC by sending prepare messages to participants and participants either accept or reject the prepare messages. In the second phase, coordinators send out commit messages if all participants accept the prepare messages and abort messages, otherwise. Both coordinators and participants write Write-Ahead Log (WAL) to persist volatile states, such that the commit decisions for recovery purposes.

3. DESIGN

DTranx adopts the SEDA architecture to reach the optimal performance in each server and achieves serializability by combining OCC and 2PC protocols. Furthermore, it introduces a NVM-based WAL design with a garbage collection mechanism to effectively and efficiently reclaim logs.

3.1 Architecture Overview

As shown in Figure 2, DTranx follows the SEDA design and invents three categories of stages: Service, Internal, and Daemon. Service stages handle Remote Procedure Calls (RPCs). For example, ClientService accepts transaction commit requests from clients and TranzService processes 2PC requests from peer servers. Internal stages manage local shared resources. For example, LockService maintains locking states and WAL writes logs to persistent storage. Daemon stages run background tasks. For example, GC periodically reclaims logs and TranxAck sends commit results from coordinators to participants.

To further exploit concurrency in SEDA, DTranx adjusts each of the stage components. First, DTranx removes the dynamic control of the thread pools but statically assigns thread numbers for each stage, after which threads are bound to physical CPU cores. We found out that one thread for each stage yields better performance when context switching is rare than that of multiple threads for each stage when context switching happens frequently. However, dynamic control of the thread pools is enabled in certain stages where handlers might be blocking. For example, Storage launches multiple threads to handle I/O requests which involves blocking system calls. Besides the core bindings for DTranx threads, kernel Interrupt Request (IRQ) threads are bound to CPU cores as well since I/O throughput is severely affected otherwise.

Second, DTranx adopts lock free queues as the incoming event queues such that the enqueue/dequeue operations on the queues are nonblocking and it achieves high throughput...
without compromising consistency. The lock free queues utilize atomic primitives to reserve a spot and then proceed to read/write in non-critical sections. In addition, multiple queues are created in each lock free queue to spread loads.

Third, DTranx reduces queue element construction and destruction costs by pushing element pointers, instead of the element itself into the lock free queues and allocating an element pool to store destructed elements. For example, Service stages get elements from the pool when new requests come and Internal stages put elements to the pool when requests are completed.

3.2 Serializability

DTranx combines OCC and 2PC protocols to guarantee serializability, following Alexander’s hybrid OCC scheme that embedded lock acquisition and validation in the 2PC. The main benefit compared to distributed Two-Phase Locking(2PL) is that locks are only held during the commit time and DTranx employs parallel validation for better scalability. The detailed protocol flows are shown in Figure 3. Additionally, if all data items in a transaction are stored in the same server, 2PC is converted to One-Phase Commit(1PC) to reduce latencies.

Initially (Stage 1), transactions read data without locks and clients keep track of the read items, read item versions, and write items in the local buffer. At commit time (Stage 2), clients choose a server as the coordinator and send it the transactions. During the first phase, coordinators initiate 2PC by first sending prepare messages to participants. Then, participants lock both the read and write items, check the read item versions, write WAL logs and reply to the coordinator. During the second phase, coordinators wait for responses from all participants, then decides whether to commit or abort, and notifies all participants of the agreed result. However, if any participant aborts in the first phase, the coordinator immediately sends out abort messages without waiting for all replies. Finally, participants write WAL logs, update database states, and unlock all relevant data.

**Proof of Serializability**

**Assumption:** Two phase locking(2PL) ensures serializability, see proof at [12].

**Method:** We reduce the hybrid OCC to 2PL. Using action abbreviations L (Locking), C (Checking), U (Unlock), R (Read), W (Write) and object abbreviations r (read items), w (write items). Concatenated action and object symbols represent tasks, e.g., Lr means “lock read items”. The sequencing abbreviation “...” binds two actions and enforces an “execute before” local order and → binds two tasks and enforces an “execute before” distributed order. Our transactions can thus be represented as R → Lw-Cr → W-Urw. The Cr action validates the read items. If any read item has been changed after it was read, the transaction aborts, releasing all locks. If not, our successful transaction is equivalent to Lr-R → Lw-Cr → W-Urw, thus Lr-R → Lw → W-Urw. In this way, all locking actions precede all unlocking actions, which is 2PL. Unlocking after committing to the database avoids cascading rollbacks. For successful transactions, the serialization point is the moment when all the write locks are granted.

**Deadlock** Common deadlock avoidance methods are timeout, wait-for graph, ordered locking and timestamps with wait-die or wait-wound mechanisms. SiloR avoids deadlocks by enforcing a global order on the locking sequences, necessitating multiple round trips in distributed environments. The wait-for graph introduces too much network traffic and timestamps method requires a global synchronized clock, which will become the bottleneck or single point of failure. Our deadlock avoidance method aborts transactions immediately if read locks are not granted and waits for a configurable time period (e.g. 50ms) before aborting write lock requests. However, if the data is currently exclusively locked, the write lock request is aborted immediately. If a transaction is aborted since write locks are not granted, DTranx retries committing it after an exponential timeout. If a transaction is aborted since read locks are not granted, DTranx restarts it immediately. This is because read lock request denial indicates there are concurrent transactions updating the same item and retrying committing will fail again.

**Livelock** Write starvation rarely happens since write lock requests are blocked for a short fixed period and exponential backoff technique is adopted to reduce the probability of lock conflicts when the transactions are retried. The same goes for read starvation since the number of read items are usually much larger than write items.

3.3 Persistent Memory

In distributed systems, the logging module plays a critical role in failure recovery. WAL is the persistent copy of the volatile states that are subject to failures due to power outage and kernel hanging. However, persisting WAL to the durable storage results in long latencies. With the advances of the NVM technologies, the performance gap between in-memory and persistent storage accesses is narrowing. Thus, we propose a WAL design based on NVM and introduce a garbage collection mechanism to effectively and efficiently reclaim the limited NVM space.

3.3.1 Log Design

The logging module is designed in three vertical layers: NVM library, LogManager and TranxLog. The NVM library provides the basic interface to persistent memory to create files, read and write data. LogManager structures the log into a list of log files, calculates checksums, and sup-
ports block read/write operations. Lastly, TranxLog offers high level abstractions for distributed transaction logs and presents a continuous and append-only log.

We use Intel's NVM [1] library to manipulate memory mapping for log files in persistent memory. After log files are mapped to the memory space, writes are immediately durable after being flushed from cache to memory(e.g. using clflush in the x86 instruction set). Two adjustments of the NVM library are made. First, a read pointer is added to the original NVM library to provide an on-demand read interface. Second, internal write locks are disabled since only one thread is launched in the WAL stage, thus no race conditions.

On top of NVM library, LogManager organizes the logs in a list structure such that logs of variable sizes are supported. To reduce I/O system calls and reach higher throughput, reads and writes are block based and logs in the same block are buffered in memory. In addition, checksums are calculated and written for each block to detect data corruption. TranxLog serializes transaction logs such as CommitLog that records commit states for coordinators and ReadyLog that records ready states for participants. Then, TranxLog separates WALs into files whose names are set to their creation timestamps. Thus, the file with the smallest timestamp is the oldest one, with which the garbage collector starts. On the other hand, reclaiming old log files does not interfere with current transactions since current transactions are appending logs to new log files.

### 3.3.2 Garbage Collection

Since WAL is written as transactions are committed, its size would increase indefinitely if DTranx does not reclaim the WAL of complete transactions. WAL for transactions that reach consensus are not required during recovery. Therefore, we introduce a state transition based garbage collection mechanism to identify unnecessary logs without performance hiccups. In particular, each transaction is assigned with a unique TranxID that combines the ServerID and a local monotonically increasing 64-bit integer. Since ServerIDs are assigned as the server indexes in the group membership store in the replicated state machine Raft, each TranxID is guaranteed to be distinct among servers. Moreover, each server keeps updating the largest committed TranxID, LC_TranxID, where transactions with TranxIDs less than LC_TranxID have reached consensus. Then, each server broadcasts its LC_TranxID and stores the LC_TranxID from other servers in a fixed-size GC log. The benefits of the GC log are twofold: it is fixed size space usage and it enables WAL reclamation.

The state transition flow is illustrated in Figure 4. On the one hand, each transaction has a state to represent the current stages in the 2PC and each server has a GC state to record completed transactions where LC_TranxIDs are calculated. On the other hand, there are volatile and non-volatile states where the nonvolatile states are durable copies of the volatile ones. For example, WALs are the nonvolatile copies of 2PC states including Start, Prepare, Ready, Commit, and Abort. GCLog is the nonvolatile state persisting the GC state. Although both WALs and GCLog are persistent copies of the volatile states, their orders of updating volatile and nonvolatile states differ. For WALs, nonvolatile 2PC states are updated after WALs are written in order for the transactions to be recoverable. For the GCLog, GC state is updated before GCLog for two reasons. First, the history can be replayed as long as WALs are not reclaimed yet. Second, accumulating in-memory GC states and writing to the GCLog in batch is more I/O efficient. For coordinators, GCThread periodically collects the volatile GC states and updates the GCLog, after which WALs containing only completed transactions are claimed and the LC_TranxIDs are broadcasted to all the other servers. For the participants, GCBroadcast thread passively receives the broadcasted LC_TranxID, updates the local GC state and GCLog, and then claims WALs.

Not only does the state transition help to reclaim WALs, it is also utilized to clean the aborted transaction IDs in the lock service, which are referenced to avoid faulty re-lock situations. For example, after a participant receives the prepare request of transactionA and its volatile state is checked to be Start, an abort request of transactionA arrives, changing the volatile state to Abort. It is possible that abort requests come before prepare requests are done since the coordinator immediately sends out abort requests if any participants aborts. Note that these two requests are processed concur-

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**Figure 4:** Coordinator and Participant state transition. Black arrows show the state transitions and red arrows show the ordered steps of garbage collection.
rently. Then, the prepare requests lock the data items and these locks will never be unlocked. Nonetheless, the committed transaction IDs are not stored in the lock service since coordinators only send out commit requests after all participants agree to commit, in which case it is impossible that prepare and commit requests are processed concurrently.

4. IMPLEMENTATION

In this section, we explain the implementation details that optimize DTranx performance.

4.1 Cache

DTranx enables client side cache to avoid excessive network traffic. The caching policy works as follows: (1) Data cache is updated if read or commit requests succeed. (2) Data cache is invalidated if commit requests fail. In addition, DTranx servers piggyback the updated data in the response to failed commit requests such that the clients can update the local cache and read requests in the retrying transaction can read from the local cache.

On the other hand, DTranx enables server side database cache to serve read requests in lower latency and it adopts the write-through strategy for durability.

4.2 Exactly-Once RPC

There are three different RPC calls corresponding to the three stages in Figure 2: Read requests from clients to servers; Commit requests from clients to servers; and, Transaction requests from servers to servers. Duplicate processing of RPCs would lead to system failures in certain cases. For example, if DTranx servers process a duplicate prepare request after the corresponding commit request is done, the locking service would lock the data items and future transactions would not be able to update these data. Therefore, DTranx should guarantee to process each RPC exactly once.

First, we guarantee at least once delivery by resending messages on the sender side if no responses are received within a timeout. We build the RPC protocol based on the ZeroMQ library, which automatically resends messages if they are lost. In addition, DTranx implements the retrying mechanism itself when no responses are received since at least once delivery in ZeroMQ does not indicate at least once delivery in DTranx. For example, if servers are restarted after the ZeroMQ library receives a message but before the DTranx system detects the message, ZeroMQ does not retry the message and the message is lost.

Second, we guarantee at most once processing by blocking duplicate messages. We assign distinct IDs for each RPC message and receivers record the IDs of completed messages. Read requests are never blocked since they are idempotent. For Commit requests, each message has a clientID and messageID where the clientID is distinct for each TCP connection and the messageID is monotonically increasing for each client. ClientIDs are assigned by the ZeroMQ library when the connection is established. For Transaction requests, each message has a unique transaction ID (TranxID) and a message type. TranxID is the concatenation of the distinct coordinator server ID and a monotonically increasing integer. And, there are four message types corresponding to the four Transaction requests in Figure 3: Prepare, Ready, Commit, and Abort. With at least once delivery and at most once processing, each message is processed exactly once.

4.3 LevelDB

We choose levelDB as the local database implementation since it is lightweight and efficient compared to multi-version KV stores. To validate the read items during the OCC commit, DTranx keeps a version number for each key value pair by storing the combination of the real value and a version number as the value in levelDB. The real value and the version number are separated by a special delimiter, such as #. When clients send read requests, servers interpret the values retrieved from levelDB and returns the value and the version number to the clients. When clients send Commit requests, servers increment the corresponding version numbers by 1 if transactions commit.

4.4 Fault Recovery

As the cluster size increases, the probability of server failures will increase considerably. For example, if the aggregated MTBF (Mean Time Between Failure) of a server is 1 year including disk failures, network failures etc., then in a cluster of 100 servers, there is a server failure every 3 to 4 days on average. DTranx triggers the recovery process in two stages: local recovery and global recovery. Local recovery reapplies local logs by updating databases if transactions commit and lock data items if no agreement has been reached. In addition, DTranx fills the TranxID space with aborted transactions. Missing transactions are possible when servers crash immediately after read item checking fails in coordinators. Global recovery repairs transactions of which commit results can not be decided unilaterally. It is initiated after local recovery to inquire transaction states from other involved servers. Specially, if the coordinator is in Prepare state and all participants are in Ready states, neither committing nor aborting violates distributed consensus. DTranx chooses to abort them such that the clients can assume the transaction failure if no responses are received.

On the other hand, DTranx starts service stages in Figure 2 after local recovery such that the changes from completed transactions are applied and in-memory states of ongoing transactions are stored. However, the order between service stage startup and global recovery does not matter and DTranx chooses to start service stages before global recovery to reduce the service downtime.

4.5 Optimization

In order to achieve better performance, multiple optimization techniques are applied. The most significant techniques are listed below.

- **Delayed In-Memory Reclamation** DTranx reclaims the volatile Commit/Abort states in participants when the servers are under light loads to avoid performance hiccups.
- **Batch Ack Phase** DTranx delays the second phase (Ack phase) of 2PC when coordinators send transaction commit results. We delegate the Ack phase to a separate stage, TranxAck in Figure 2, to reduce the transaction latency and offload the high processing demand of the ClientService stage.
- **Core Bindings** We manually analyze the queue size for each stage and bind the threads to physical cores in an optimal way. The best core binding strategy yields almost 6 times higher throughput than the worst. In
the future, we plan to explore how to automate the core bindings to attain the best performance based on the number of CPU cores available.

5. EVALUATION

Our benchmark tests are run on a Cloudlab cluster with 36 machines. Each of the machines has Intel E5-2660 v3, 20 2.6GHz cores, equipped with 130GB RAM, 480GB solid state disk at 6GB/sec, and 10Gbps Ethernet card. We emulate the NVM by enabling DAX support in Linux to create a PM-aware environment. The DRAM based emulation is adopted since current persistent memory latency is comparable to DRAM and NVM was not available. For example, STT-RAM achieves 10ns write latency compared to 50ns DRAM latency. NVM throughput is also far beyond the current usage as shown in Figure 7. To generate workloads, we use Yahoo Cloud Serving Benchmark (YCSB) C++ version and add DTranx and HyperDex support. YCSB clients are running in separate servers from the cloud servers that accommodate the DTranx system. For each test result, the average of 3 runs are reported.

5.1 Environment Setup

First, we evaluate DTranx with a database of 120 million key value items in a 36-node cluster. Test data keys are generated as integers from 1 to 120 million and values are 100 bytes of random characters. Transactions are categorized into read and update transactions. Read transactions only contain read items and update transactions contain 1 write item. The total number of read/write items in one single transaction is uniformly distributed between 1 and 3.

5.2 Transaction

In Figure 6, DTranx is compared with Hyperdex Warp that supports distributed transactions. Only successful transactions are counted in the throughput metric. DTranx shows approximately 30% higher throughput than Hyperdex and DTranx degrades slowly as the percentage of update transactions increases. Moreover, DTranx maintains high commit success rates. For example, DTranx reaches 99.65% success rate for 50% read workloads. On the other hand, Hyperdex shows high throughput but the software is unstable and periodically fails due to internal assertion errors, leading to low success rates. For example, several servers crashed during the 95% read workload, causing 58.96% success rates and 275.72k ops/sec throughput. To remedy the crash effect, we restarted the servers manually after each run. There are three reasons why DTranx outperforms Hyperdex. First, DTranx follows the highly concurrent SEDA architecture with lock free queues and stages are bound to physical cores, utilizing all CPU power and avoiding context switching overhead. Second, DTranx integrates the NVM based log that bypasses system calls like sync/fsync, reducing log persistence latencies. Third, DTranx applies various optimization techniques, such as an allocated element pool, batch ack phase, and optimal core binding strategy. Furthermore, strace reveals that Hyperdex does not synchronize data to physical storage devices immediately after write log calls. While the Hyperdex paper supports fault recovery by replication, that version of the software is not publicly available. Lastly, the average latency for DTranx is below 2ms when the throughput is 50% of the maximum and it increases to 10ms when the throughput reaches the maximum.

5.3 Scalability

In this experiment, scalability tests are run against cluster of 3, 9, 18 and 36 servers. Corresponding to the cluster size, 10, 30, 60, 120 million keys are inserted into DTranx. As shown in Figure 7, the throughput shows linear increases as more nodes are involved. For example, with pure read workloads, throughput reaches 574.76k reqs/sec with 36 nodes. In addition, workloads with various mixture of read and update transactions are benchmarked. Even with 50% read workloads and 50% update workloads, the throughput is 60% to 85% of that with pure read workloads. The high scalability of DTranx results from our efficient hybrid commit protocol design that minimizes the critical section of distributed locking and reduces the 2PC to 1PC whenever possible.

5.4 Persistent Memory

Two experiments are conducted to validate the effectiveness and efficiency of the NVM based log. Both experiments are run with 36 servers and 95% read workloads. In Figure 8, the instant throughput is plotted with and without GC.
The GC process doesn’t affect normal transactions when WALs are GC’ed every 10 seconds since the reclamation of volatile states that affects normal transactions is delayed until servers are in light loads. The system with SSD log shows 19k ops/sec on average, which is 30 times slower than that with NVM log. In Figure 7a, we measure the space over time with and without GC to show the GC efficiency. The logs are NVM files of 100MB size so that the space usage changes in units of 100MB. The GC mechanism successfully keeps log space usage low since DTranx reclaims the transactions from WALs much faster than it writes them. After 120 seconds, tests are complete and the log usage with GC converges to 200 MB (one for GLog and one for the current WAL).

6. RELATED WORK

DTranx is a highly concurrent and transactional KV store that integrates various techniques from concurrent programming, database, and NVM fields. DTranx combines OCC and 2PC to move the locks to the commit time and employ parallel validation for better scalability. Experiments show that DTranx offers higher throughput than the state-of-the-art system, Hyperdex, and DTranx displays high scalability for various workloads.

7. CONCLUSIONS

We propose a transactional and scalable key value store that utilizes non-volatile memory based log with an effective and efficient garbage collection mechanism. To exploit the multi-core machines, we adapt the SEDA architecture with lock free queues and apply an optimal core binding strategy. Moreover, DTranx combines OCC and 2PC to move the locks to the commit time and employ parallel validation for better scalability. Experiments show that DTranx offers higher throughput than the state-of-the-art system, Hyperdex, and DTranx displays high scalability for various workloads.

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