Virtual Disk Snapshot Management at Scale

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Abstract

Contrary to the other resources such as CPU, memory and network, for which virtualization is efficiently achieved through direct access, disk virtualization is peculiar. In this paper we make four contributions. Our first contribution is the characterization of disk utilization in a public large scale cloud infrastructure. It reveals the presence of long snapshot chains, composed sometimes of up to 1000 files. Our second contribution is to show by experimental measurements that long chains lead to performance and memory footprint scalability issues. Our third contribution is the extension of both the Qcow2 format and its driver in Qemu to address the identified scalability challenges. Our fourth contribution is the thorough evaluation of our prototype, called SQUEMU, demonstrating that it brings significant performance enhancements and memory footprint reduction. For example, it improves the throughput of RocksDB by about 48% compared to vanilla Qemu on a snapshot chain of length 500. The memory overhead on that chain is also reduced by 15x.

1 Introduction

Virtualization is the keystone technology making cloud computing possible and therefore enabling its success. However, virtualization, and thus cloud computing, comes at the cost of a certain overhead on application performance. That overhead have been well studied [1, 3, 9, 10, 14, 20, 24, 33]. Although it concerns all types of resources (CPU, RAM, network, disk), they are not all affected with the same intensity. Figure 1 shows the performance degradation coming from virtualization for a wide range of benchmarks including Stream [11] (memory intensive), NPB [6] (CPU intensive), netperf [28] (network intensive), as well as the Linux dd command (disk intensive, throughput-oriented) and fio [16] (disk intensive, latency-oriented), when they run in AWS EC2 (t2.medium instance type), Microsoft Azure (Standard_B2s instance type), a virtualized private cloud, and on bare metal private cloud without virtualization1. We use the latter as the baseline. We can observe that the two disk-intensive applications (dd and fio) experience the highest slowdown. For fio, it is about 1,639× the degradation experienced by NPB.

Surprisingly, contrary to the other resource types, very few research work focuses on improving storage virtualization in the cloud. It is important to fill this gap, in particular in the context of the explosion of data-centric application’s (big data, ML and AI trends) popularity. Disk virtualization is peculiar as it is still implemented through complex multi-layered architectures [15, 36]. Another illustration of the singularity of disk virtualization is the fact that it is generally achieved through the use of complex virtual disk formats (Qcow2, QED, FVD, VDI, VMDK, VHD, EBS, etc.) that not only perform the task of multiplexing the physical disk, but also need to support standard features such as snapshots/rollbacks, compression, and encryption. These indirections are the source of the disk virtualization overheads.

This paper focuses on Linux-KVM/Qemu (hereafter LKQ), a very popular virtualization stack. LKQ supports several virtual disk formats, among which Qcow2 [27] is widely adopted in production [21]. Our cloud partner, which is a large scale public cloud provider with several datacenters spread over the world, relies on LKQ and Qcow2. A salient feature provided by Qcow2 is the capacity to create incremental Copy-On-Write (COW) snapshots (backing files) in order to save the state of the virtual disk at a given point in time and to reduce storage space usage. The virtual disk of a VM can thus be seen as a chain linking multiple backing files. In this paper, we identify and solve virtualization scalability issues on such snapshot chains.

Our first contribution (§3) is the characterization of disk usage in the infrastructure of our cloud partner. We found that snapshot operations are very frequent in the cloud (some VMs are subject to more than one snapshot creation per day) for three main reasons. First, cloud users leverage snapshots to periodically create recovery points for fault tolerance reasons. Second, cloud users and providers use snapshots to achieve efficient virtual disk copy operations, as well as to share some elements such as the OS/distribution base image between several distinct virtual disks. Third, cloud providers use the snapshot feature to transparently distribute a virtual disk, made of multiple chained backing files, among several storage servers, in effect going above the boundaries of a single physical server. This is achieved for load balancing reasons, and also tackles resource fragmentation issues. For all these reasons, we observed that

1We chose t2.medium and Standard_B2s to match the VM size that we used in our private cloud.
the length of a chain can be very high. We identified chains composed of up to 1,000 backing files. To our knowledge, this is the first paper performing such characterization. Prior works [2, 5, 29] mainly focused on the characterization of virtualized CPU and memory utilization in the cloud.

Our second contribution (§4) is to show by experimental measurements that long chains pose both performance and memory footprint scalability issues. For illustration, using a synthetic benchmark based on dd, we measured up to 91% of IO throughput decrease and up to 180 × memory footprint increase for a chain composed of 1,000 backing files. We found that the origin of this problem lies in the fundamental design of the Qcow2 format: the fact that the Qcow2 driver in Qemu manages each backing file individually in a recursive fashion, without a global view of the entire chain composing the virtual disk. The analysis of other formats (such as FVD [30], see §7) shows that they use a similar approach.

Our third contribution (§5) is to address these scalability issues by evolving Qcow2 and introducing two key principles: 1) direct access upon an I/O request, regardless of their position in the chain; 2) the use of a single Qcow2 metadata cache, avoiding memory duplication by being independent of the chain length. The implementation of these principles raises three challenges. First, we should allow backward compatibility, which is necessary to facilitate the adoption of our solution by cloud operators. Second, we should preserve all Qcow2 features including compression, encryption, etc. Third, important optimizations such as prefetching that come naturally with the current Qcow2 format should also be preserved. To cope with the above challenges, we slightly extend the Qcow2 format in order to indicate, for each cluster of the virtual disk, the backing file it is contained in. To preserve backward compatibility, we rely on reserved bits in Qcow2’s metadata. We implement these principles by extending on the one hand the Qemu’s Qcow2 driver and the snapshot operation on the other hand. We thoroughly evaluate our prototype in several situations: various disk sizes, chain lengths, cache sizes, and benchmarks. Our solution tackles Qcow2’s scalability issues regarding IO performance and memory footprint. For example, on a virtual disk backed up by a chain of 500 snapshots, RocksDB’s throughput is increased by 48% versus vanilla Qemu. The memory overhead on that chain is also reduced by 15 ×.

Overall we make the following contributions:

- We characterize for the first time virtual disk management and usage in a large scale could provider.
- We assess for the first time performance/memory footprint scalability issues in Linux-KVM/Qemu/Qcow2, a popular virtualization stack, and explain the origin of the problems.
- We introduce two principles for addressing this problem.
- We implement these principles in Qemu while preserving all its features.
- We evaluate our prototype in various situations, demonstrating the effectiveness of our approach.

The rest of the paper is organized as follows. §2 presents the background. §3 presents virtual disk characterization results. §4 presents and assesses the scalability issues handled in this paper. §5 presents our design to addresses the identified scalability issues. §6 presents the evaluation results of our design. §7 presents the related work. §8 concludes the paper.

2 Linux-KVM/Qemu/Qcow2

Our goal is twofold: the characterization of virtual disk management in a public large scale cloud and handling of two scalability issues happening in long snapshot chains. This section presents the necessary background to understand our contributions.

Qcow2 Overview. The Qcow2 format enables copy-on-write snapshots by using an indexing mechanism implemented in the format and managed at runtime in the Qcow2 driver, running in Qemu, to map guest IO requests addressing virtual sectors/blocks to host offsets in the Qcow2 file(s). An overview of the Qcow2 format is given on Figure 2. Without any snapshot, a virtual disk is contained in a single file. The file is divided into units named clusters, that can contain either metadata (e.g., a header, indexation tables, etc.) or data that represent ranges of consecutive sectors. The default cluster size is 64 KB. Indexation is made through a 2-level table, organized as a radix tree: the first-level table (L1) is small and contiguous in the file, while the second-level table (L2) could be spread among multiple non-contiguous clusters. The header occupies cluster 0 at offset 0 in the file and the L1 tables come right after the header. For performance reasons, L1 and L2 entries are cached in RAM (see bellow).

Qcow2 Snapshotting. A Qcow2 virtual disk file F can be linked to a backing file, i.e. a file that will be queried for clusters that are not present in F. Today, the most common way to create a live incremental snapshot of a virtual disk F for a given VM is to create a new empty Qcow2 file E and set it as the current disk (called active volume) for the VM while the previous virtual disk F is set as the backing file for E. In that way, all write operations made by the VM will be directed to the active volume (E) while read operations will be directed either to E if the addressed sectors are present there, or to backing files if not. With time, backing file chains can become very long (see §3).

Qcow2 Cache Organization. To speed up access to L1 and L2 tables, Qemu caches them in RAM. It creates and manages one cache for the active volume and one cache per backing file. Each cache is managed independently from the others. In the following, we describe how the cache works. Qemu
maintains a separate cache for the L1 table and a cache for L2 tables entries. With its small size, the entire content of L1 is loaded in RAM at VM boot time. The cache of L2 entries is populated on-demand, with a prefetching policy. We therefore focus on the caching of L2 entries as they are likely to suffer from misses, thus influence IO performance. When there is a cache miss, Qemu brings into the cache a set of L2 entries, a slice of configurable size, among which the entry at the origin of the miss. The slice is also the granularity of the cache eviction policy, which is LRU. A cache entry includes: the file offset of the slice (noted \( l_{2\_slice\_offset} \)), the number of threads which currently uses the slice (noted \( \text{ref} \)), the actual L2 entries composing the slice, and a field indicating whether a data cluster referenced by a L2 entry has been modified (noted \( \text{dirty} \)).

Qcow2 Cache Utilization. Every IO request issued by the guest OS to virtual disk \( vb \) traps inside Qemu. It is then handled by a thread running the para-virtualized disk driver in Qemu. One of its main goals is to translate \( vb \) to a data cluster offset inside the active volume or a backing file.

From \( vb \), Qemu computes \( l_{2\_slice\_offset} \), \( l_{2\_slice\_index} \), and \( l_{2\_index} \). Having this, Qemu looks if there is an entry in the cache that matches \( l_{2\_slice\_offset} \). If it exists, then Qemu increments the corresponding \( \text{ref} \). Next, thanks to \( l_{2\_slice\_index} \), it reads the L2 entry. If the latter describes an allocated data cluster (hereafter “cache hit”), then Qemu reads the offset of the data cluster. If the cluster is not allocated (hereafter “cache hit unallocated”) then Qemu considers the cache of the next backing file in the chain. If the slice is not in that cache, then Qemu will try to fetch it from the actual backing file associated with the current cache. If the slice does exist on disk, then it is brought into the cache. Otherwise, Qemu considers the cache of the next backing file, and so forth.

Some additional actions are performed for write requests. First of all, the \( \text{dirty} \) field of the slice is set to 1. If the L2 entry is found in a backing file (not the active volume), Qemu allocates a data cluster on the active volume and performs the copy-on-write. If, despite the whole chain scanning, the L2 entry is not found, then Qemu just creates a new data cluster in the active volume. In any case, Qemu configures L1 and L2 tables accordingly, both on disk and in the active volume’s cache. A cache entry can be evicted either when the VM is terminated or when the cache is full.

IO Request Journey on a Chain. Qemu manages a chain snapshot-by-snapshot, starting from the active volume. Figure 3 illustrates the journey of an IO request for a chain of size 2: the base image (B) and the active volume (V). We assume that all L2 indexing caches are empty. Let us assume that cluster number 2 is the target cluster, and it resides in B (meaning that it has not been modified since the creation of V).

1. The driver starts by parsing V indexing cache. To handle the cache miss, Qemu performs a set of function calls with some of them accessing over the network the Qcow2 file to fetch the missed entry from V’s L2 table. According to its prefetching feature, Qemu fetches a slice of L2 table entries from V and fills V indexing cache. In Figure 3, we assume that the size of a slice is 2 entries. Thus, V’s cache includes at the end of the first cache miss handling process two valid entries: cluster 1 and cluster 2. After this step, pv driver hits V’s cache, but the state of cluster 2 is marked unallocated because the references data cluster resides on B.

Figure 3: The journey of an IO request.
hit unallocated event triggers the same Qemu functions used for handling a cache miss. For cache hit unallocated events, Qemu moves to the parent snapshot (B). In fact, at VM startup, Qemu initializes a linked list corresponding to the snapshot chain of the VM’s virtual disk. The caches of all the snapshots are also initialized at that time. 6 The first access to B’s cache generates a miss 7. After handling this miss (8–10), the offset of cluster 2 is returned to the driver. From there, 11 the latter can issue the IO request (consider read here).

3 Virtual Disk Management Characterization

This section presents a characterization of virtual disk management by our cloud industry partner. We explore various metrics, namely VM requested storage size, snapshot chain length, snapshot chain sharing, and snapshot creation frequency. The study targets a datacenter located in Europe. The number of VMs booted in 2020 in this region is 2.8 millions which corresponds to one VM booted every 12 seconds, demonstrating the large scale of our study. The software and workloads running in the VMs are highly varied. However it is worth noting that our partner specializes in business-to-business, hence the VMs run enterprise workloads as opposed to private individual ones. Similar to existing cloud providers, the region runs VMs internal to our partner in addition to client VMs. In the region, the VMs’ virtual disks are backed up by Qcow2 chains.

Virtual Storage Resources Requested. Figure 4 shows the CDF of the virtual disk size for active volumes as well as backing files created over a day in 2020. We have separated first party resources, used by the provider to operate the cloud and to provide other services, as well as third party resources, used by final customers. We can see volumes and snapshots requesting up to 10 TB. 10 GB volumes corresponds to the default virtual disk size, and represents 30% of the first party requests in both volumes and snapshots. In the case of the third party, the most popular size among clients is around 50 GB with 40%.

Take-away 1: Virtual disks of all sizes, up to 10 TB, are used in the infrastructure. The most popular requested sizes are 10 GB (first party) and 50 GB (third party).

Chain Length. Over the entire year 2020, we performed a daily measurement of the length of each chain in the infrastructure. This covers a wide dataset, with the number of daily chains considered being in the order of the hundreds of thousands. Snapshots (backing files) can be created by clients but also by the cloud provider, for example when it is decided that a new storage node should host part of a virtual disk, or when the client creates a virtual disk as a copy of another one.

Figure 5 presents the evolution of the longest chain’s length over the period. As we can see, there is always a chain with at least a length of 800 snapshots, and the longest chain can have a length of up to more than 1,000.

We studied in details a daily measurement made during the period when the longest chain was of a length superior to 1000. Figure 6 shows the CDF for chains and files (active volumes and backing files) with respect to the chain length (for a file, to the length of the chain it belongs to). Most of the chains are relatively small: chains of length 10 or lower represent nearly 50% of the total number of files, and more than 80% of the files. This is because, for a subset of the chains, the backing file merging operation, named streaming, is triggered around size 30. That operation merges the layers corresponding to multiple backing files into a single one. The files that can be merged in this way correspond to unneeded snapshots, i.e. deleted client snapshots as well as the ones made by the provider. Streaming helps reduce the size of some chains, however note that valid (non-deleted) client snapshots cannot be merged. Further, although they are infrequent, there is a non-negligible number of chains of size 100 and above.
Figure 7: Chains can share backing files following a virtual disk copy, or when using a common base image.

Figure 8: For each chain C of a daily measurement, percentage of backing files shared with another chain according to the length of C.

**Take-away 2:** Long chains, with up to 1,000 backing files, do exist. The chain size threshold triggering streaming will cap the maximum size of many chains in the infrastructure.

**Chain Sharing.** Certain backing files are shared, and belong to chains corresponding to different virtual disks. The two main sources of sharing are virtual disk copy operations, as well as the use by multiple VMs of virtual disk base OS distribution images offered by the provider. A virtual disk copy is made by transforming the active volume into a backing file, and creating 2 new active volumes on top, forming 2 chains: all the backing files are thus shared between the 2 chains. This is illustrated on the bottom of Figure 7. Concerning base OS distribution images, they are generally themselves composed of multiple snapshots corresponding to the different construction steps followed by the provider: the corresponding backing files are thus shared between all chains using a given base image. This is illustrated on top of Figure 7.

In Figure 8, each point corresponds to a chain of the daily measurement previously considered. The chain’s length is indicated by its X value, and the number of backing files in the chain that are shared with at least another chain is indicated by the Y value. Note that in theory, a chain of length \( N \) can share from 0 up to \( N - 1 \), files with other chains, i.e. all backing files without counting the active volume. Overall, the degree of sharing is highly variable among chains. We can observe a significant amount of chains of variable length with no sharing at all (Y value of 0). The high number of chains with a length \( N < 30 \) allows us to witness, for these chains, almost all possible degrees of sharing (from 0 to \( N - 1 \)). The large number of points around size 30 corresponds to the high number of chains of that length, due as explained above to the streaming threshold being set to 30. Although the number of chains of size superior to 30 is smaller, one can still observe a variable degree of sharing for some of these. Note that, base OS images are generally made of around 5 chained backing files, so most of the sharing presented in Figure 7 is due to virtual disk copies.

**Take-away 3:** Backing files can be shared between several chains when multiple VMs use the same base OS image, or to achieve virtual disk copy. The degree of sharing among chains in the infrastructure is highly variable. Past a certain length (5+), most of the sharing is due to virtual disk copies.

**Snapshot Creation Frequency.** Finally, we investigated the frequency of snapshot (i.e. backing files) creation. We looked in our daily measurement, for each snapshot creation operation, the time elapsed since the creation of the previous link in the chain (either a backing file, or the active volume for a first snapshot).

This data is presented on Figure 9. Each point corresponds to a set of snapshot creation operations, placed on the Y axis into buckets corresponding to different elapsed time windows since the last link creation. Each point’s X value corresponds to the position in the chain of the created backing files. Finally, the size and color of each point denotes how many snapshot creation operations are represented by the point, as a percentage of the total number of operations counted in our daily measurement.

As one can observe, the majority of the snapshots are made on chains of size inferior to 30. This is due to the high number of chains of these sizes, stemming from the streaming threshold set to 30. Further, although the frequency of snapshot creation is overall highly variable, an important number of snapshots are created with a relatively high frequency (daily
or more). Past work [25] noted peaks at up to 58 snapshots per hour. One can also observe that the long chains are the result of relatively frequent (daily/weekly) snapshotting done by clients (i.e. non-mergeable through streaming).

**Take-away 4:** Although the snapshot creating frequency varies widely among chains, a non-negligible amount of chains experience high frequency snapshotting. Long chains belong to this subset, with daily/weekly snapshot created. These snapshots are made by clients and cannot be merged with streaming.

4 Problem with Long Snapshot Chains

4.1 Origins of Long Snapshot Chains

The emergence of long snapshot chains in modern virtualized environments is due to a combination of factors. First, for data backup/fault tolerance purposes, most cloud providers offer to the client the possibility to create disk snapshots either on a regular basis, for example every 24 hours, or on-demand through an API. The chain length will thus grow according to the snapshot frequency. Even in the case the client deletes certain snapshots, they are kept by the provider as they form a necessary part of the Qcow2 chain backing the disk the VM in question is currently using. Second, snapshots may be performed by the cloud provider itself due to thin provisioning strategies: virtual disk space being allocated on-demand, a disk may grow above the boundaries of the physical disk storing it and, combined with distributed storage, a snapshot allows to have the virtual disk transparently continue to grow on another physical disk without data transfer. Although they are not visible by the client, such snapshots will be placed in the chains in the same way as the client-made snapshots and will participate to the chain’s size increase.

An intuitive way to tackle the long chains issue is to merge snapshots that have been deleted and as such compact the chain, in a process referred to as streaming. These techniques are quite limited as the cloud provider has no control over the client-made snapshots. Furthermore, streaming seriously impacts guest I/O performance: we measured the disk latency from the guest with ioping on a standard SSD (WD Blue) and noted a 100x increase during streaming. Streaming can be quite long according to the size of the merged snapshots, and a streaming operation needs to abort in case the client decides to reboot/halt the VM.

**Take-away 5:** Long chains are due to the client- as well as provider-made snapshots, and to the limitations of the methods (e.g., streaming) to reduce their length.

4.2 Problem Statement

From the illustration presented in Figure 3, the reader can intuitively see the two scalability issues posed by Qcow2 for long chains. The first one is memory footprint increase, caused by L2 entry duplication in indexing caches. In Figure 3, cluster 1 and cluster 2 are present in the two indexing caches. The second consequence is the negative impact on I/O request latency. We can formalize the average cache miss cost (Y) using this equation:

\[
Y = (Hit_{\text{guest}} \times T_M) + (Miss_{\text{guest}} \times (T_D + T_L + T_F)) + (UnAllocated_{\text{guest}} \times T_F) \times N
\]

where \(T_M\) is the RAM access time (about 100ns), \(T_D\) is the disk access time (about 80\mu s), \(T_L\) is the time to traverse all software and network layers (about 1\mu s), and \(N\) is the chain length, \(Hit_{\text{guest}}\), \(Miss_{\text{guest}}\), and \(UnAllocated_{\text{guest}}\) are respectively the hit, miss and unallocated events ratios. According to the fact that \(T_D, T_L, T_F\) are too high compared to \(T_M\), even a small miss and unallocated ratio will lead to significant performance degradation [30]. This degradation is exacerbated for long snapshot chains.

4.3 Assessment

A VM running on a long Qcow2 snapshot chain sees its performance and memory footprint seriously impacted. To demonstrate these points, Figure 10 shows the evolution of these two metrics for a VM running on a virtual disk with variable chain sizes, ranging from 0 to 300 snapshots. The total virtual disk size is 20 GB and each snapshot contains an incremental layer of 60 MB. All files reside locally on the host’s SSD. The VM has 4 GB of allocated RAM, 4 vCPUs and runs Ubuntu 18.04. The read throughput is measured within the VM by reading the entire disk with dd right after 1) a first call to dd on the entire disk to ensure L1/L2 caches are fully populated and 2) a guest page cache drop to assure that the Qcow2 file is accessed. The memory footprint is measured from the host as the hypervisor’s peak Resident Set Size (RSS) observed during the execution of the dd command.

As one can observe, although with small chains the read throughput is not substantially impacted, when the chain size grows this metric drops significantly. On a virtual disk with a chain size of 300, the read throughput only reaches 39 % of what can be achieved on a disk with no snapshots. Regarding memory consumption, with no or a few snapshots the memory overhead that Qemu presents on top of the 4 GB used by the VM is negligible. However, with long snapshot chains that overhead becomes significant: with 300 snapshots, 711 MB of additional RAM are consumed by Qemu. Third, we used

![Figure 10: I/O performance and memory footprint evolution with snapshot chain size.](image-url)
the massif heap profiler of Valgrind to investigate memory consumption during the dd test on the 300 snapshots-long experiments, and discovered that the memory footprint increase is due to various data structures that are allocated on a per-snapshot basis. The main culprit for the high memory consumption with long chains is the L2 indexing cache. There is one Qcow2 driver instance running in the hypervisor for each Qcow2 snapshot in a chain. Although the maximum L2 cache size defaults to 1 MB [8], in our experiment we set it to 2.5 MB which is enough to manage a 20 GB disk – setting it lower seriously impacts performance. However, because there is one cache per driver instance and one instance per snapshot, one can conclude that the cache-related memory footprint increases linearly with the number of snapshots in the chain. These numbers were gathered on Qemu 4.2 but we also confirmed this behavior on the latest (v6.0) version. We focus on 4.2 in the rest of this paper as it is the version used by our cloud provider partner.

We also profiled the Qemu hypervisor from the host during the execution of the aforementioned dd test on the 300 snapshots-long case and found that the guest only executes for 7% of the time. Qemu’s disk driver threads consume the remaining time.

**Take-away 6:** Long chains lead to memory footprint and IO performance scalability issues.

5 SQUEMU: Scalable Qemu

This section presents a new version of both Qemu and Qcow2 which tackles the two scalability issues identified in the previous section, regarding performance and memory consumption. Ideally, both metrics should be as independent as possible from the length of the backing file chain length.

5.1 Principles and Challenges

SQUEMU relies on two key principles, illustrated on Figure 11: 1) direct access to on-disk indexing/data clusters, regardless of their position in the chain, upon guest I/O requests; 2) the use of a single unified indexing cache, avoiding cache entries duplication by being independent of the chain length. In the rest of the document, we note vQemu and vQcow2 respectively vanilla Qemu and Qcow2 current format. We apply the first principle through a slight but backward-compatible modification of vQcow2, requiring the storage of additional metadata in virtual disk images, as well as an update to the Qcow2 driver in the Qemu storage stack: we call that evolution SQUEMU for scalable Qemu. Meanwhile, applying the second principle only requires a carefully modification of the Qcow2 driver.

A major challenge the implementation of SQUEMU faces regards its transparent and fast integration within the infrastructure of our cloud partner (and within cloud infrastructures in general). Our solution should first be compatible with the different backends that can hold disk backing files in today’s cloud infrastructure: these can be stored directly on the host disk but also accessed by the host through the network and served by centralized NFS servers or distributed file systems. Hence we propose to modify a popular existing disk format rather than propose a new one [30]. A related challenge is also backward compatibility: existing Qcow2 images lacking our format’s metadata should still work with our updated version of Qemu (without performance/memory consumption gains on long chains), and images using our format should also work with vanilla version of Qemu that do not run our updated Qcow2 driver (once again without gains on long chains). Alternatively, vanilla disk images can be easily converted to our format to benefit from the performance/memory footprint enhancement on long chains.

5.2 Format Improvement

When a guest issues an IO request, vQemu sequentially scans the active volume and all the backing files in the chain until the proper one is found, which is not efficient. We propose to slightly update the Qcow2 format as well as its management algorithms in Qemu in order to eliminate that chain scanning operation. To this end, we introduce a new metadata in the format indicating, for each data cluster, the backing file that contains the latest (i.e. valid) version of the cluster. We call this metadata the *backing_file_index*. We leverage unused bits in L2 table entries to do so. We use 16 bits to encode *backing_file_index* in each L2 entry.

5.3 Unified Cache and Direct Access

With direct-access, we maintain a single unified cache for the entire disk, independently of the length of the backing file chain. Our cache has the same organization as the vanilla Qcow2 cache presented in Section 2. As a reminder, a cache entry corresponds to a slice and contains: l2_slice_offset (playing the role of the tag), ref, dirty, and the L2 entries composing the slice. As noted in the previous section, in SQUEMU a L2 entry contains *backing_file_index* in addition to the default Qcow2 values.

Contrary to the vanilla version where l2_slice_offset was specific to each backing file, in our version, l2_slice_offset is related to the active volume. In addition, one can find in the same slice, L2 entries describing data clusters belonging to distinct backing files. Therefore, the read and write operations are performed as follows in SQUEMU. Let us consider vb the offset of a virtual block that the guest wishes to read. Using the same functions a vQemu,
SQEMU computes l2_slice_offset, l2_slice_index and l2_index. If both the slice and the L2 entry exist in the unified cache and that backing_file_index contained in the L2 entry corresponds to the active volume, then there is a cache hit and the offset of the cluster data to be read is in the L2 entry. If backing_file_index does not correspond to the active volume, this is a cache hit unallocated. SQEMU locates on disk the backing file corresponding to backing_file_index and reads from it the slice at offset l2_slice_offset. Let s_p be that slice and s_v be the slice currently contained in the unified cache. SQEMU traverses all s_p entries and updates the L2 entries in s_v with the corresponding contents in s_p under the following condition: the value of backing_file_index of the L2 entry in s_v is lower or equal to that of backing_file_index of the L2 entry in s_p. We call “cache correction” these replacement operations. Then it sets dirty to 1 in s_v, so that the slice will be written to disk when it is evicted from the cache. If the L2 entry does not exist in the slice there is a cache miss and the entry needs to be allocated as in VQemu. This means that the guest is asking for a data cluster which does not yet exist on the virtual disk. If the slice is not yet present in cache, there is a cache miss and the slice is either fetched from the active volume if it exists, or allocated if not. These operations are similar to vQemu.

5.4 Snapshotting

In vQemu, when a snapshot is created, a new Qcow2 active volume is created, with very few information (the header, the L1 table and refcounts). We update the snapshot creation logic to copy to the newly created active volume the entire content of both L1 and L2 tables from the previous active volume, now a backing file. The algorithm that we implement is as follows. Let new_volume be the file that will become the new active volume and old_volume the old one. We intervene at the creation of any Qcow2 file. Let new_l1 be the new L1 table and old_l1 the L1 table of old_volume. After the allocation of new_l1, we parse all the old_l1 entries. For each entry we create the corresponding L2 table in new_volume, then we set the current new_l1 entry with the offset of that L2 table. Let new_l2 be that new L2 table and old_l2 be the L2 table pointed to by old_l1 in old_volume. Then we copy the whole content of old_l2 to new_l2.

As a consequence, a new active volume always contains all L2 tables of the previous backing files. The copy of L2 tables may lengthen disk snapshotting time compared to the vanilla version. We could have implemented a copy on-demand solution, however that would mean impacting the critical path of I/O requests. This approach would increase tail latency as it requires chain scanning to find the valid backing file. The evaluation results show that the disturbance brought by the snapshot operation upon guest I/O performance is largely acceptable, as the total size of L2 tables is in the order of MB. In addition, we think that VM owners are likely to accept the small price of a slight increase in snapshotting time, to benefit from an important boost in I/O performance.

6 Evaluation

Here we present an evaluation of SQEMU, aiming to answer the following three questions:

Q1) Does SQEMU eliminate the memory footprint scalability issue of VQemu? (§6.2)

Q2) Does SQEMU eliminate the IO performance scalability issue of VQemu? (§6.3-6.4)

Q3) To what extent SQEMU increases snapshotting time and disk overhead? (§6.5)

6.1 Evaluation Setup

Methodology. We systematically compare SQEMU with VQEMU. We evaluate several configurations by varying three parameters: the chain length (1-1,000); the virtual disk size (50GB, 150GB); as well as the cache size (from 30% to 100% of the cache size needed to hold the entirety of L2 entries to index a full disk, i.e. from 1.9 MB to 6.25 MB for a 50GB disk size, and from 5.6 MB to 18.75 MB for 150 GB). For all experiments, valid clusters are uniformly distributed on the backing files of the disk’s chain. The virtual disk is populated at 90% with random data for experiments with micro-benchmarks using the Linux dd command, and at 25% for experiments with macro-benchmarks using the RocksDB client [7]. The release of SQEMU includes a highly configurable chain generation script.

Otherwise indicated, the size of the L2 cache is set so that it can hold all L2 entries to index the entire disk. All results presented in this section are an average value of 5 runs.

Testbed. To have a representative test environment, we employ 2 servers, one being the compute node running VMs, and the other the storage node holding virtual disk files. Each server is equipped with a 32 cores Intel Xeon Gold CPU, clocked at 2.10 GHz, 192 GB of RAM, Samsung MZ7KM480HMHQ0D3 SATA SSD. They are linked with a 10Gbps Ethernet connection. The storage node serves the virtual disk files through NFS. Both servers run Debian 10 with Linux 4.19.0 as host OS. All VMs run Ubuntu 18.04 with Linux 4.15.0 and are configured with 4GB of memory, 4 vCPUs. Otherwise indicated, the virtual disk size is 50GB.

Metrics and Benchmarks. We collect two kind of metrics, high-level and low-level metrics. The former are those which directly impact the end-user perceived Quality of Service. We consider VM startup time, memory overhead, application execution time and, I/O disk throughput. The memory overhead is the additional memory consumed by Qemu on top of the VM’s allocated pseudo physical memory. Low-level metrics represent internal costs that help explaining high-level metrics. They are: the total number of cache misses, the total number of cache hit unallocated, and the cache lookup latency. The lookup latency is the time taken to find the valid offset of
a data cluster in the caching system. Storage benchmarks are run in the guests. We use microbenchmarks, including Linux `dd` (which sequentially read the entire disk from the guest i.e. `dd if=/dev/sda of=/dev/null bs=4M`) and fio [16], as well as macrobenchmarks, RocksDB-YCSB [7] and a measurement of the VM boot time.

### 6.2 (Q1) Memory overhead

For this experiment we measured Qemu’s resident set size after having read the entire disk from the guest using `dd`, and subtracted from this measurement the amount of RAM given to the VM (4GB) to compute Qemu’s overhead. Figure 12 shows the results. One can observe that sQEMU brings a significant reduction of the memory overhead when the chain length increases. The memory savings are as follows: 205 MB for a chain length of 50 (3.9× reduction), 2303 MB for a length of 500 (15.2×), and 4289 MB for a length of 1,000 (17.6×). Although it scales much better than vanilla Qemu, sQEMU’s memory overhead still slightly increases with the chain size. This is due to other per-snapshot data structures in Qemu that are not directly related to the caches. Finally, note that sQEMU comes at the cost of a slight memory footprint increase over vanilla when the disk has no or a very small number of snapshots — a cost that is amortized by the better scalability starting from 5 snapshots.

### 6.3 (Q2) Low-level Metrics

We use the same setup as in the previous section.

**Cache Misses and Cache Hit Unallocated.** We instrumented sQEMU and vanilla Qemu to measure the number of cache misses, the number of cache hits unallocated, and the number of caches accesses per backing file of the chain. Figure 13 shows the results.

We can see that sQEMU leads to less cache misses compared to vQEMU, see Figure 13a. We measure up to 10× for chain length 1,000. This difference is explained by the fact that vQEMU does not implement a cache correction mechanism as we do in sQEMU (see §5). Therefore, when an L2 entry is only present in the cache of the backing file of index m in the chain, vQEMU will generate n−m+1 cache misses walking the chain to get it, where n is the chain length.

Concerning the number of cache hits unallocated, it is constant under sQEMU, see Figure 13b. The increase in comparison to a virtual disk composed of a single active volume (chain length 1) is less that 1% for the chain length 1,000. Concerning vQEMU, the number of cache hit unallocated increases 10,000,000× for the chain length 1,000. This is once again explained by the fact that vQEMU looks up several caches during the chain walk.

In the experiment with a chain of length 500, we count the total number of cache lookups and plot their distribution, according to which backing file in the chain holds the requested data, on Figure 13c. As expected caches are much more accessed under vQEMU compared to sQEMU, due to the chain walks. The gap is about 1,500%. The spike that appears for backing file zero, the base virtual disk image, corresponds to boot of the VM. In fact, during that time, several IO read requests are performed on read only files (such as vmlinuz). The spike on snapshot 500 corresponds to the accessed made on the active volume.

**Cache Lookup Latency.** We measured the cache lookup latency on two chain lengths: 1 and 100. Figure 14 presents the distribution of cache lookup latencies for all IO requests performed during the execution of the `dd` benchmark. We can observe that for both systems, the mean latency value changes according to the chain length. However, sQEMU leads to a...
better latency compared to vQEMU when the chain length increases: the mean latency is 490 ms under vQEMU and 270 ms with sQEMU, i.e. 1.8x faster. Contrary to vQEMU, latency values under sQEMU are located around two mean values 120 ms and 270 ms. 120 ms corresponds to the cache hit mean latency while 270 ms corresponds to the cache hit unallocated mean latency. Note that theoretically, according to the direct access principle implemented by sQEMU, only one value of cache hit unallocated latency can be observed compared to vQEMU. We do not observe the same kind of distribution under vQEMU because in this experiment, data clusters are uniformly distributed over all backing files. Therefore, most IO operations lead to a variable amount of cache hits unallocated, i.e. chain walks of variable length, according to the target data location in the chain. This translates into highly variable and on average higher latencies in vQEMU.

6.4 (Q2) High-level Metrics
6.4.1 Micro-benchmarks

Disk Throughput: Linux dd. The throughput of dd is presented in Figure 15 for both systems managing chains of various sizes. We can observe no degradation under sQEMU while vQEMU severely degrades the throughput of dd when the number of backing files increases. vQEMU incurs a slowdown of up to 84% for the chain length 1,000.

Impact of the Cache Size with fio. We studied the effect of varying the cache size for sQEMU and vQEMU. In this experiment we use of chain of length 500 and set the total cache size used by vQEMU to be equal to that used by sQEMU. Because vQEMU uses one cache per layer in the chain, when sQEMU is given a cache size of S, vQEMU would get S/L with L being the chain length. We vary the cache size given to each system from 1MB to 4GB, and measure the disk read throughput with fio performing random reads of small size (4 KB) on the disk node in /dev.

Figure 16 shows the results. We can observe that sQEMU significantly outperforms vQEMU in all cases. With both systems, performance are sensitive to the cache size. Concerning vQEMU, performance steadily increase up to 4 GB of cache. This is due to the large amount of memory required by this multi-caches solution. Regarding sQEMU, although peak performance are also achieved at 4 GB (6 MB/s vs. 2.5 MB/s for 1 MB of cache), from 32 MB the payback from adding more cache size diminishes significantly. This value thus represents for a good trade-off between near-peak performance and memory footprint. This demonstrates the high efficiency of sQEMU vs. vQEMU.

6.4.2 Macro-benchmarks

VM Boot Time. VM boot time is a critical metric in the cloud [18, 22]. Figure 17 compares the time it takes to boot a VM under sQEMU and vQEMU while varying the chain length and the virtual disk size. The boot time increases rapidly with the chain length under vQEMU: it goes from about 10 seconds on a chain of size 1 to more than 40 seconds (4×) on a chain of size 1000. On the contrary, with sQEMU that increase is moderate: from 10 seconds to 17 seconds (1.7×).

The increase in boot time for sQEMU can be explained by the slight increase of the number of cache misses and cache hit unallocated discussed above. We can see that the size of the virtual disk does not really influence the results.

Cloud Workload: RocksDB-YCSB. We created a RocksDB database that fills 40% of the VM disk size, and populated using the YCSB client, generating a uniform distribution of valid clusters of the Qcow2 chains generated. We use YCSB-C, which simulates a user performing read-only requests. We experimented two L2 cache sizes (1 MB and 3
6.5 \((Q_3)\) Overhead

As stated in §5.4, when creating a snapshot under sQEMU, L2 tables are copied to the new created file. This may incur two overhead types: disk usage and snapshotting time.

**Disk space.** The disk space overhead per snapshot depends on both VM’s disk size and cluster size, as well as the number of allocated clusters in the disk. We can model that overhead in the worst case scenario, i.e. when every cluster is allocated (the disk is full), as follows: given \(S_{SQ} \) and \(S_{VQ} \) being respectively the size of a newly created (i.e. empty) snapshot under sQEMU and vQEMU, we compute the disk size of \(S_{SQ} \) using the following formula:

\[
S_{SQ} = S_{VQ} + \frac{VM\_disk\_size \times cluster\_size}{L2\_entry\_size} \quad (2)
\]

By default, an L2 entry is 8 B, a cluster is 64 KB, and \(S_{SQ} \) is 256 KB. Using the above formula, we can compute \(S_{SQ} \) while varying the VM disk size from 50 GB to 200 GB. That per-snapshot overhead is shown in Figure 19a. It increases linearly with the size of the VM disk. To compute the total disk overhead (still in the worst case), the per-snapshot cost needs to be multiplied by the chain length. Recall that from our characterization, we observed that the dominant virtual disk size in the cloud is 50 GB, giving according to our model a per-snapshot overhead is about 6 MB. This gives a total overhead, in the worst case, of 60 MB for a chain of length 10 (0.1% of the virtual disk size), 600 MB for length 100 (1.2%), and 6,000 MB for length 1000 (12%).

**Snapshotting Time.** We measured the time spent to create a new snapshot under sQEMU and vQEMU for different VM disk sizes. The results are presented on Figure 19b. Due to the copy of all L2 entries, sQEMU takes much more time to create a snapshot compared to vQEMU. For a 50GB VM, we need about 70ms to create a snapshot under sQEMU and 7× less time under vQEMU. Furthermore, this overhead increases with the VM disk size. Indeed, for a 200GB VM, the snapshot creation time under sQEMU is about 12× that of vQEMU. Nonetheless, in absolute the snapshot creation latency is quite low under sQEMU (in the order of ms). It allows for a relatively high snapshot frequency.

7 Related Work

The systems software literature contains relatively few contributions regarding storage virtualization. We identified a few recurrent topics, presented below.
Virtual Disk Formats. The research topic that is the most relevant to our work relates to proposals of new and more efficient virtual disk formats [19, 30, 32].

Fast Virtual Disk [30] is a virtual disk format proposed by IBM in 2011, that advocate for high flexibility with many configurable options and increases I/O performance by avoiding the use of a host filesystem, reducing the size of on-disk metadata, and using an on-disk journal. FVD supports only internal snapshots, which means that all the chain is stored in a single file. This may not be as flexible as the external snapshots offered by the format we focus on, QCOW2, for example when subsets of a chain need to be stored on different storage nodes for load-balancing or capacity reasons. It is also unclear how FVD performs on long chains composed of hundreds or thousands of snapshots. The system we propose is an evolution of QCOW2 which is backwards compatible with vanilla QCOW2 disk, something that makes adoption much easier versus proposing an entirely new format. Finally, FVD can be considered as depreciated as it we developed for Qemu 0.14, dating from 2011, and has not been ported to modern versions.

Parallax [19] is a distributed architecture storing virtual disk images that allows the use of commodity servers as storage backends, as opposed to high-end storage arrays/switches. Among other features, Parallax offers low-overhead and high-frequency snapshots and note, similar to our work, that the performance overhead and memory consumption of traditional formats such as QCOW2 increases with snapshot chain sizes. Similar to FVD, migrating an existing cloud environment to Parallax would require significant changes to the virtualized storage system’s architecture, whereas we rely on the widely used QCOW2 format, and are backward-compatible with environments that do not use our system. Further, contrary to our QCOW2 format, Parallax does not support sharing of virtual disk images, a feature heavily used in the industry to lower storage overheads of commonly used volumes such as base images.

Storage Performance and Availability during VM Migration. Noting that VM migration significantly disrupts guest I/O performance, a few papers [13, 35] focus on maintaining good storage performance and availability during migration. Netchannel [13] proposes various techniques to maintain local/remote virtual disk availability during migration. One is the ability to seamlessly switch the physical device associated with a virtual one. Another proposed technique is the capacity for migrated VMs initially plugged to a local disk on the host to transparently keep using that disk through a proxy once they are migrated to another host. In another study [35], the author propose to study the storage I/O behavior of guests to infer the most efficient data transfer schedule to reduce disruption as much as possible during VM migration.

Scheduling Impact on Virtual Storage Performance. Several studies [12, 23] noted that VM scheduling could have a non-negligible impact on guest I/O performance. The authors of a study [23] characterize the impact on processor and I/O performance of various VM scheduler configurations, for concurrently-running guest with CPU- and bandwidth-intensive, as well as latency sensitive behaviors. In another paper [12], the authors propose a guest task-level priority boosting technique to selectively increase the priority of I/O-bound task to increase storage performance while maintaining CPU fairness.

Virtualized Storage Performance and Power Consumption. Other studies focus more generally on virtualized storage performance and power consumption [31, 34]. Ye and al. [34] note that existing consumption reduction reduction techniques focusing on non-virtualized HDDs do not apply well in a virtualized setting. They propose to bridge the semantic gap between VM and VMM through several techniques tailored for such environments, reducing disk spin-ups and increasing disk sleep times, in order to save energy. Another paper [31] focuses on the particular problem of interrupt delivery to VMs, including the ones coming from block devices. The authors propose an optimized interrupt delivery system for KVM. It is mostly evaluate on network workloads but also shows moderate performance improvements on storage workloads.

Cloud Storage and File Systems. Finally, several papers [4, 17, 26] focus on cloud storage and filesystems. The Frugal Cloud File System [26] proposes integrating multiple services (AWS EBS, Azure Cache, etc.) into a single solution that aims to be flexible from the performance and costs point of views. DepotSky [4] introduces a cloud-based storage system targeting security/dependability by spreading and replicating storage over multiple clouds. Depot [17] proposes a cloud storage system that can tolerate buggy clients and servers in order to minimize trust assumptions.

8 Conclusion

We present, for the first time, the characterization of virtual disk management in a large scale public cloud using the Linux-KVM-Qemu/QCOW2 virtualization stack. Among other results, our analysis revealed the presence of long snapshot chains, leading to scalability issues for both memory footprint and performance. We present SQEMU, a solution to these two issues, in the form of a slight extension of the Qcow2 format while preserving backward compatibility. We built SQEMU following the principles of direct access and single indexing cache, regardless the chain length. We evaluate SQEMU extensively compare it with vanilla Qemu using a wide range of benchmarks, demonstrating that our solution effectively tackles the above issues. For instance, SQEMU improves the I/O throughput of RocksDB by up to 48% compared to VQEMU, and reduce the memory footprint by 15x, when the chain length is 500.

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