Approximating Dynamic Time Warping Distance Between Run-Length Encoded Strings

Zoe Xi
Massachusetts Institute of Technology, Cambridge, MA, USA

William Kuszmaul
Massachusetts Institute of Technology, Cambridge, MA, USA

Abstract
Dynamic Time Warping (DTW) is a widely used similarity measure for comparing strings that encode time series data, with applications to areas including bioinformatics, signature verification, and speech recognition. The standard dynamic-programming algorithm for DTW takes $O(n^2)$ time, and there are conditional lower bounds showing that no algorithm can do substantially better.

In many applications, however, the strings $x$ and $y$ may contain long runs of repeated letters, meaning that they can be compressed using run-length encoding. A natural question is whether the DTW-distance between these compressed strings can be computed efficiently in terms of the lengths $k$ and $\ell$ of the compressed strings. Recent work has shown how to achieve $O(k\ell^2 + \ell k^2)$ time, leaving open the question of whether a near-quadratic $\tilde{O}(k\ell)$-time algorithm might exist.

We show that, if a small approximation loss is permitted, then a near-quadratic time algorithm is indeed possible: our algorithm computes a $(1 + \epsilon)$-approximation for $\text{DTW}(x, y)$ in $\tilde{O}(k\ell/\epsilon^3)$ time, where $k$ and $\ell$ are the number of runs in $x$ and $y$. Our algorithm allows for DTW to be computed over any metric space $(\Sigma, \delta)$ in which distances are $O(\log n)$-bit integers. Surprisingly, the algorithm also works even if $\delta$ does not induce a metric space on $\Sigma$ (e.g., $\delta$ need not satisfy the triangle inequality).

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1 Introduction

Dynamic Time Warping (DTW) distance is a well-known similarity measure for comparing strings that represent time-series data. DTW distance was first introduced by Vintsyuk in 1968 [40], who applied it to the problem of speech discrimination. In the decades since, DTW has become one of the most widely used similarity heuristics for comparing time series [28] in applications such as bioinformatics, signature verification, and speech recognition [20, 34, 33, 1, 14, 43].
Approximating DTW Distance Between RLE Strings

Consider any two strings $x$ and $y$, with characters taken from some metric space $(\Sigma, \delta)$. For example, in many applications, we have that $\Sigma = \mathbb{R}^c$ for some parameter $c$ and that $\delta(a, b) = \|a - b\|_2$ computes $\ell_2$ distance. Define a time warp of $x$ (and similarly of $y$) to be any string $x'$ that can be obtained by warping the letters in $x$, where warping a letter means replacing it with $\geq 1$ consecutive copies of itself. The DTW-distance $DTW(x, y)$ is defined to be

$$\min_{|x'| = |y'|} \sum_{i=1}^{|x'|} \delta(x'_i, y'_i),$$

where $x'$ and $y'$ range over all time warps of $x$ and $y$.

The most fundamental question concerning DTW is how to compute it efficiently. Vintsyuk showed that, given strings $x$ and $y$ of length $n$, it is possible to compute $DTW(x, y)$ in $O(n^2)$ time [40]. His algorithm, which was one of the earliest uses of dynamic programming, continues to be taught in textbooks and algorithms courses today.

For many decades, it was an open question whether any algorithm could achieve a running time of $O(n^{2-O(1)})$. (Interestingly, it is known that one can shave small sub-polynomial factors off of the running time [22].) A major breakthrough occurred in 2015, when Abboud, Backurs, and Williams [2] and Bringmann and Künnemann [12] established conditional lower bounds prohibiting any strongly subquadratic-time algorithm for DTW, unless the Strong Exponential Time Hypothesis (SETH) fails.

This lower bound puts us in an interesting situation. On one hand, the classic $O(n^2)$-time algorithm is often too slow for practical applications. On the other hand, we have good reason to believe that it is nearly optimal. This has led researchers to focus on forms of beyond-worst-case analysis when studying the DTW problem.

An especially appealing question [39, 21] is what happens if $x$ and $y$ both contain long runs of repeated letters. In this case, the strings can be compressed using run-length encoding (RLE). For example, the string “aaaaabbc” has RLE encoding “(a, 5), (b, 2), (c, 1)”. If a string $x$ has $k$ runs, then it is said to have an RLE representation of length $k$.

It is known that, if $x$ and $y$ each contain $k$ runs, then $DTW(x, y)$ can be computed in $O(k^3)$ time [21].\(^1\) It is still an open question whether it is possible to significantly reduce this cubic running time, and in particular, whether a near-quadratic time algorithm might be possible.

This paper: A Near-Quadratic Approximation Algorithm

We show that, if a small approximation loss is permitted, then a near-quadratic time algorithm is indeed possible. Consider any two run-length encoded strings $x \in \Sigma^n$ and $y \in \Sigma^m$, where $x$ has $k$ runs and $y$ has $\ell$ runs. Let $\delta$ be an arbitrary distance function $\delta : \Sigma \times \Sigma \rightarrow [\poly(n)]$ mapping pairs of characters to $O(\log n)$-bit nonnegative integers. (Perhaps surprisingly, our algorithms will not require $\delta$ to satisfy the triangle inequality, or even to be symmetric.)

Our main result is an algorithm that computes a $(1 + \epsilon)$-approximation for $DTW(x, y)$ in $\tilde{O}(k\ell/\epsilon^2)$ time\(^2\). In the special case where $\Sigma$ is over Hamming space (i.e., $\delta(a, b)$ is either 0 or 1 for all $a, b \in \Sigma$), the running time of our algorithm further improves to $\tilde{O}(k\ell/\epsilon^2)$.

\(^1\) More generally, if $x$ contains $k$ runs and $y$ contains $\ell$ runs, then the time becomes $O(k\ell^2 + k\ell^2)$.

\(^2\) Here we are using soft-O notation to mean that $\tilde{O}(k\ell/\epsilon^2)$ is equivalent to $O((k\ell/\epsilon^2)\polylog(n))$. 

Our algorithm takes a classical geometric interpretation of DTW in terms of paths through a grid, and shows how to decompose each path in such a way that its components can be efficiently approximated. This allows for us to reduce the problem of approximating DTW-distance between RLE strings to the problem of computing pairwise distance in a small directed acyclic graph.

Other related work

In addition to work on run-length-encoded strings [39, 21], there has been a recent push to study other theoretical facets of the DTW problem. This includes work on approximation algorithms [25, 3, 42], low-distance-regime algorithms [25], communication complexity [11], slightly-subquadratic algorithms [22], reductions to other similarity measures [25, 37, 36], binary DTW [26, 38], etc.

All of these results (along with the results in this paper) can be viewed as part of a larger effort to close the gap between what is known about DTW and what is known about its closely related cousin edit distance, which measures the number of insertions, deletions, and substitutions of characters needed to turn one string $x$ into another string $y$. Like DTW, edit distance can be computed in $O(n^2)$ time using dynamic programming [40, 35] (and can be computed in slightly subquadratic time using lookup-table techniques [31]). Also like DTW, edit distance has conditional lower bounds [12, 2, 25] prohibiting strongly subquadratic time algorithms.

When it comes to beyond-worst-case analysis, however, edit distance has yielded much stronger results than DTW: it is known how to compute a constant-approximation for edit distance in strongly subquadratic time [5, 24, 10, 15, 4, 9, 6, 17]; it is known how to compute the edit distance between RLE strings in $\tilde{O}(k\ell)$ time [19, 8, 18, 29, 13, 7, 32, 23, 30]; and if two strings $x$ and $y$ have small edit distance $k$, it is known how to compute the edit distance in $O(|x| + |y| + k^2)$ time [16, 27].

Whether or not any of these results can be replicated for DTW remains the central open question in modern theoretical work on DTW. There are several reasons to believe that DTW computation should be more challenging than edit distance. Whereas edit distance satisfies the triangle inequality, DTW does not (for example, if we take $\Sigma = \{0, 1\}$, then $DTW(111110, 100000) = 0$, $DTW(100000, 000000) = 1$, and $DTW(111110, 000000) = 5$). This erratic behavior of DTW seems to make it especially difficult to approximate. Additionally, whereas almost all work on edit distance focuses on insertion/deletion/substitution costs of 1, work on DTW must consider arbitrary cost functions $\delta$ for comparing characters. Finally, it is known that the problem of computing edit distance actually reduces to that of computing DTW [25], indicating that the latter problem is at least as hard (although, interestingly, this reduction does not apply in the run-length encoded setting).

Our paper represents the first evidence that an $\tilde{O}(k\ell)$-time algorithm for DTW may be within reach. Such an algorithm would finally unify edit distance and DTW in the run-length-encoded setting.

2 Technical Overview

This section gives a technical overview of how we approximate DTW-distance between run-length encoded strings. To simplify exposition, we focus here only on the big ideas in the algorithm design and we defer the detailed analysis to later sections.
Throughout the section, we consider two strings $x$ and $y$ of length $n$ whose characters are taken from a set $\Sigma$ with a symmetric distance function $\delta : \Sigma \times \Sigma \rightarrow \mathbb{N} \cup \{0\}$. Our only assumption on $\delta$ is that $\delta(a, b) \in \{0, 1, 2, \ldots, \text{poly}(n)\}$ for all $a, b \in \Sigma$. (We do not need the triangle inequality on $\delta$.) Let $k$ and $\ell$ be the number of runs in $x$ and $y$, respectively. We will describe a $(1 + O(\epsilon))$-approximate algorithm that takes $\tilde{O}(k\ell/\text{poly}(\epsilon))$ time.

How to think about DTW

There are several mathematically equivalent ways (see, e.g., [25, 22, 21]) to define the dynamic time warping distance between $x$ and $y$. In this paper, we work with the geometric interpretation: consider an $n \times n$ grid where cell $(i, j)$ has cost $\delta(x_i, y_j)$; consider the paths through the grid that travel from $(1, 1)$ to $(n, n)$ via steps of the form $(1, 0)$ (a horizontal step (h-step) to the right), $(0, 1)$ (a vertical step (v-step) up), and $(1, 1)$ (a diagonal step (d-step) to the upper right); the cost of such a path is the sum of the costs of the cells that it encounters, and $DTW(x, y)$ is defined to be the smallest cost of any such path. For an example, see Figure 1, which shows an optimal full path for computing $DTW(aaabbbddd, aabcd)$ = 1, where the $\delta$-function measures the distance between characters in the alphabet.

Note that the $i$-th column of the grid corresponds to $x_i$ and the $j$-th row of the grid corresponds to $y_j$. Thus, each run $x_{i_0}, \ldots, x_{i_1}$ in $x$ corresponds to a sequence of adjacent columns $i_0, \ldots, i_1$ in the grid, and each run $y_{j_0}, \ldots, y_{j_1}$ in $y$ corresponds to a sequence of adjacent rows $j_0, \ldots, j_1$ in the grid.

If we want to design an algorithm that approximates $DTW(x, y)$ in $\tilde{O}(k\ell/\text{poly}(\epsilon))$ time, then it is natural to think about the grid as follows. We break the grid into blocks by drawing a vertical line between every pair of runs in $x$ and a horizontal line between every pair of runs in $y$; and label the blocks $\{B_{i,j}\}_{i \in [k], j \in [\ell]}$, where block $B_{i,j}$ corresponds horizontally to the $i$-th run in $x$ and vertically to the $j$-th run in $y$. All of the cells within a given block $B$ have the same cost, which we refer to as $\delta(B)$. We may also use $\delta_{i,j}$ for $\delta(B_{i,j})$. We refer to the first/last row of each block as a lower/upper horizontal boundary and to the first/last column of each block as a left/right vertical boundary.

Finally, it will be helpful to talk about sequences of blocks that are adjacent horizontally or vertically. An $h$-block segment consists of a sequence of consecutive blocks lined up horizontally. Formally, given $i_1 \leq i_2$ in $[k]$ and $j$ in $[\ell]$, we use $B_{i_1, i_2, j}$ for the $h$-block segment $B_{i_1, j}, B_{i_1+1, j}, \ldots, B_{i_2, j}$. Similarly, a $v$-block segment consists of a sequence of consecutive blocks lined up vertically – we use $B_{i, j_1, j_2}$ for the $v$-block segment $B_{i, j_1, j_1+1}, \ldots, B_{i, j_2}$. In the same way that we can talk about the four boundaries of a block, we can talk about the four boundaries of a given $h$-block or $v$-block segment.

Intuitively, since there are $\Omega(k\ell)$ blocks, our goal is to design an algorithm that runs in time roughly proportional to the number of blocks.

How to think about the optimal path

Let $P$ be a minimum-cost path through the grid. We can decompose the path into a sequence of disjoint components $P_1, P_2, P_3, \ldots$, where each component takes one of two forms:

1. A horizontal-to-vertical (h-to-v) component connects a cell on the lower boundary of some $v$-block segment to another cell on the right boundary of the same $v$-block segment.
2. A vertical-to-horizontal (v-to-h) component connects a cell on the left boundary of some $h$-block segment to another cell on the upper boundary of the same $h$-block segment.

The components $P_1, P_2, P_3, \ldots$ are defined such that the end cell of each $P_r$ connects to the the start cell of each $P_{r+1}$ via a single step (either horizontal, vertical, or diagonal).
We will now describe a series of simplifications that we can make to $P$ while increasing its total cost by at most a $(1 + O(\epsilon))$-factor. The simplifications are central to the design of our algorithm.

**Simplification 1: Rounding each component to start and end on “snap points”**

Let us call a grid cell $(i, j)$ an **intersection point** if it lies in the intersection of a horizontal boundary and a vertical boundary. (Each block contains at most four intersection points.) We call a grid cell a **snap point** if either it is an intersection point, or it is of the form $(i + (1 + \epsilon)^j, j)$ on the upper boundary of a block $B$, or it is of the form $(i + (1 + \epsilon)^j, j)$ on the lower boundary of a block $B$, or it is of the form $(i, j + (1 + \epsilon)^j)$ on the right boundary of a block $B$, or it is of the form $(i, j + (1 + \epsilon)^j)$ on the left boundary of a block $B$, where $(i, j)$ is an intersection point of the block $B$ and $t$ is nonnegative integer (since this is a technical overview, we ignore floor and ceiling issues). For each boundary cell $p$ in the grid, define $\text{snap}(p)$ to be the nearest snap point to the right of $p$, if $p$ is on a horizontal boundary, and to be the nearest snap point above $p$, if $p$ is on a vertical boundary. If $p$ is on both a horizontal boundary and a vertical boundary, then $p$ is an intersection point, so $\text{snap}(p) = p$.

How much would the cost of $P$ increase if we required each of its components to start and end on snap points? Suppose, in particular, that we replace each component $P_r$ with a component $P_r'$ whose start point $p_r$ has been replaced with $\text{snap}(p_r)$ and whose end point $q_r$ has been replaced with $\text{snap}(q_r)$. It may be that $\text{snap}(q_r)$ does not connect to $\text{snap}(p_{r+1})$, meaning that $P'_r$ and $P'_{r+1}$ do not connect properly. If this happens, however, then one can simply modify the starting-point of $P'_{r+1}$ in order to connect it to $P'_r$ (and it turns out this only makes $P'_{r+1}$ cheaper).

Let $P'$ be the concatenation of $P'_1, P'_2, P'_3, \ldots$. To bound the cost of $P'$, we can argue that the cost of each $P'_r$ is at most $(1 + \epsilon)$ times that of $P_r$. To transform $P_r$ into $P'_r$, the first step is to round the start point $p_r$ of $P_r$ to $\text{snap}(p_r)$ — one can readily see that this only decreases (or leaves unchanged) the cost of $P_r$. The second step is to round the end point $q_r$ of $P_r$ to $\text{snap}(q_r)$. For simplicity, assume that $P_r$ is an h-to-v component that starts on the lower boundary of some block $B_{i,j}$ and finishes on the right boundary of some block $B_{i,j}$. Let $(u, v)$ be the lower-right intersection point of $B_{i,j}$ and suppose that $P_r$ finishes in cell $(u, v + s)$. Then $P_r$ incurs cost at least $(s + 1) \cdot \delta_{1,j}$ in block $B_{i,j}$. Moreover, the snap point $\text{snap}(q_r) = \text{snap}(u, v + s)$ is guaranteed to be in the set $\{(u, v + s + t) \mid t \in \{0, 1, \ldots, s\}\}$. Thus the cost of traveling from $q_r$ to $\text{snap}(q_r)$ is at most $\epsilon \cdot s \cdot \delta_{1,j}$. So the cost of $P'_r$ is at most $(1 + \epsilon)$ times that of $P_r$.

By analyzing each component in this way, we can argue that $\text{cost}(P') \leq (1 + \epsilon) \text{cost}(P)$. Throughout the rest of the section, we will assume that $P$ has been replaced with $P'$, meaning that each component starts and ends with a snap point.

**Simplification 2: Understanding the structure of each component**

Next we observe that each individual component can be assumed to have a relatively simple structure. For simplicity, let us focus on an h-to-v component $P_r$ in a v-block segment $B_{i,[j,j]}$. We may assume without loss of generality that all of $P_r$’s h-steps occur together on the lower boundary of some block; and that all of $P_r$’s v-steps occur at the end of $P_r$. In other words, $P_r$ is of the form $D_1 \oplus H \oplus D_2 \oplus U$ where $\oplus$ is for path concatenation, $D_1$ consists of d-steps, $H$ consists of h-steps (along a lower boundary), $D_2$ again consists of d-steps, and $U$ consists of v-steps (along a right boundary).3 (See Figure 2 where the path $p_1 q_1 q_2 p_2 p_3$ in solid lines is such an example.)

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3 Note that the components $D_1$, $H$, $D_2$, and $U$ are each individually allowed to be length 0.
Approximating DTW Distance Between RLE Strings

Combined, these assumptions make it so that $P_r$ is fully determined by four quantities: (1) $P_r$’s start point $p_r$, (2) the block $B_{i,j}$ in which $H$ occurs, (3) the length of $H$, and (4) the length of $U$.

Define $\overline{P}_r$ to be the prefix of $P_r$ that terminates as soon as $U$ hits its first snap point. (See Figure 2 where the path $p_1q_1q_2p_2p_3^\prime$ is such a prefix of the path $p_1q_1q_2p_2p_3$.) We will see later that $\overline{P}_r$ is, in some sense, the “important” part of $P_r$ to our algorithm. Observe that $\overline{P}_r$ is fully determined by just three quantities: (1) $P_r$’s start point $p_r$, (2) the block $B_{i,j}$ in which $H$ occurs, and (3) the length of $H$.

Simplification 3: Reducing the number of options for each component

We will now argue that, if we fix the start point $p_r$, and we are willing to tolerate a $(1 + O(\epsilon))$-factor approximation loss, then we only need to consider $\text{poly}(\epsilon^{-1} \log n)$ options for $\overline{P}_r$.

We begin by considering block $B_{i,j}$ in which $H$ occurs. Let us define the sequence of blocks $B_0, B_1, B_2, \ldots$ so that $B_i = B_{i,j+i}$ and define the sequence of costs $\delta_0, \delta_1, \delta_2, \ldots$ so that $\delta_s = \delta_{i+j+s}$. We say that a block $B_s$ is extremal if $(1 + \epsilon)\delta_s \leq \delta_t$ for all $t < s$. If we are willing to tolerate a $(1 + O(\epsilon))$-factor increase in $\overline{P}_r$’s cost, then we can assume without loss of generality that $H$ occurs in an extremal block. On the other hand, there are only $O(\log^2_{1+\epsilon}(n))$ extremal blocks, so this means that we only need to consider $O(\log_{1+\epsilon}(n)) \leq \text{poly}(\epsilon^{-1} \log n)$ options for the starting point of $H$.

Next we consider the length of the horizontal sub-component $H$. If we are willing to tolerate a $(1 + O(\epsilon))$-factor increase in $\overline{P}_r$’s cost, then we can round $|H|$, the length of $H$, up to be a power of $(1 + \epsilon)$ (or to be whatever length brings us to the next vertical boundary). Thus we only need to consider $O(\log_{1+\epsilon}(n)) \leq \text{poly}(\epsilon^{-1} \log n)$ options for $|H|$.

Together, the block $B_{i,j}$ in which $H$ occurs and the length of $H$ fully determine $\overline{P}_r$. Thus, we have reached the following conclusion: if the start point $p_r$ of the component $\overline{P}_r$ is known, then there are only $\text{poly}(\epsilon^{-1} \log n)$ options that we must consider for what $\overline{P}_r$ could look like. Moreover, although we have considered only h-to-v components here, one can make a similar argument for v-to-h components.

Approximating DTW in $\tilde{O}(k\ell/\text{poly}(\epsilon))$ time

We will now construct a weighted directed acyclic graph $G = \langle V, E \rangle$ that has two special vertices $v_0$ and $v_*$, and that satisfies the following properties:
- $G$ has a total of $\tilde{O}(k\ell/\text{poly}(\epsilon))$ vertices/edges, and
- the distance from $v_0$ to $v_*$ in $G$ is a $(1 + O(\epsilon))$-approximation for $\text{DTW}(x, y)$.

This reduces the problem of approximating $\text{DTW}(x, y)$ to the problem of computing a distance in a weighted directed acyclic graph. The latter problem, of course, can be solved in linear time with dynamic programming; thus the graph $G$ gives us a $\tilde{O}(k\ell/\text{poly}(\epsilon))$-time $(1 + O(\epsilon))$-approximation algorithm for $\text{DTW}$.

We construct $G$ to capture the different ways in which path components $P_r$ can connect together (assuming that the path components take the simplified forms described above). As the vertices $v \in V$ correspond to the snap points $p$ in the grid, we can use a vertex to refer to its corresponding snap point and vice versa. We define $v_0$ to be the cell $(1, 1)$ in the grid and $v_*$ to be the cell $(n, n)$. We add edges $E$ as follows:
We connect each snap point \( p \) on a horizontal (resp. vertical) boundary to the next snap point \( q \) to its right (resp. above it).

We connect each snap point \( p \) on a right (resp. upper) boundary to any snap points \( q \) on the adjacent left (resp. lower) boundary that can be reached from \( p \) in a single step.

Each snap point \( p \in V \) has \( \text{poly}(\epsilon^{-1} \log n) \) out-edges corresponding to the \( \text{poly}(\epsilon^{-1} \log n) \) options for what a (truncated) component \( T^p \) starting at \( p \) could look like.

Note that, although we only add edges for truncated path components \( T^p \) (rather than full components \( P_\delta \)), these edges can be combined with edges of the first type in order to obtain the full component. This is why we said earlier that the truncated component is the "important" part of the component.

The paths from \( v_0 \) to \( v_\star \) in \( G \) correspond to the ways in which we can concatenate path components together to get a full path through the grid; if we assign the appropriate weights to the edges, then the cost of a path through \( G \) corresponds to the cost of the same path through the grid. The distance from \( v_0 \) to \( v_\star \) is therefore a \( (1 + O(\epsilon)) \)-approximation for \( DTW(x, y) \).

Finally, we must bound the size of \( G \). Each block contains at most four intersection points; so there are \( O(k\ell) \) total intersection points. Each intersection point creates at most \( O(\log_{1+\epsilon}(n)) \) snap points; so there are \( O(k\ell \epsilon^{-1} \log n) \) snap points (which are the vertices in \( V \)). Each snap point has an out-degree of at most \( \text{poly}(\epsilon^{-1} \log n) \). Hence we have:

\[
|E| \leq O(k\ell \epsilon^{-1} \log n) \text{ poly}(\epsilon^{-1} \log n) = \tilde{O}(k\ell/\text{poly}(\epsilon)).
\]

We can therefore compute the distance from \( v_0 \) to \( v_\star \) in \( \tilde{O}(k\ell/\text{poly}(\epsilon)) \) time, as desired.

Paper outline

For the sake of simplicity, there are a number of details that we chose to ignore in this section (such as a time-efficient construction of \( G \) and a careful proof that the modifications to \( P \) incur only a \( (1 + O(\epsilon)) \)-factor change in its cost). In the remainder of the paper, we give a formal presentation and analysis of the algorithm outlined above.

3 Preliminaries

We use \([n_1, n_2] \) for the set \{\( n_1, n_1 + 1, \ldots, n_2 \)\} consisting of all the integers between \( n_1 \) and \( n_2 \), inclusive, and use \( [n] \) as a shorthand for \([1, n] \). We use \( T_{m,n} \) for a table consisting of \( m \) columns and \( n \) rows and \( T_{m,n}[i,j] \) for the entry on the \( i \)-th column and \( j \)-th row, where \((i, j) \in [m] \times [n] \) is assumed. We may use \( T \) for \( T_{m,n} \) if \( m \) and \( n \) can be readily inferred from the context. Please note that an entry \( T_{m,n}[i,j] \) in a table should be distinguished from the value stored in the entry – when discussing the value, we shall refer to it as the content of the entry \( T_{m,n}[i,j] \).

Letters

Let us assume an alphabet \( \Sigma \), which is possibly infinite. We use \( \delta \) for a distance function on letters such that \( \delta(a, a) = 0 \) for any \( a \in \Sigma \). We do not require that \( \delta \) be symmetric or the triangular inequality \( \delta(a, c) \leq \delta(a, b) + \delta(b, c) \) hold for \( \delta \).

\[\text{Note that } G \text{ is not necessarily simple. If there are multiple ways that a component } T^p \text{ could connect two vertices } p_1 \text{ and } p_2, \text{ then there will be multiple edges from } p_1 \text{ to } p_2.\]
### Strings

We use $x$ and $y$ for strings. We write $x = (a_1, \ldots, a_m)$ for a string consisting of $m$ letters such that $x[i]$ (often written as $x_i$), the $i$-th letter in $x$, is $a_i$ for each $i \in [m]$. We use $a^n$ for a string of $n$ occurrences of $a$, which is also referred to as a run of $a$, and $\hat{x}$ for a run-length encoded (RLE) string, which consists of a sequence of runs. We use $|\hat{x}|$ and $\|\hat{x}\|$ for the length and r-length of $\hat{x}$, which are $m_1 + \cdots + m_k$ and $k$, respectively, in the case $\hat{x} = (a_1^{m_1}, \ldots, a_k^{m_k})$.

A run in a string $x$ is maximal if it is not contained in a longer run in $x$. There is a unique run-length encoding $\hat{x}$ of $x$ that consists of only maximal runs in $x$, and this encoding $\hat{x}$ is referred to as the RLE representation of $x$. We also use $\|x\|$ for the number of maximal runs in $x$ (and thus $|x| = \|\hat{x}\|$).

We use $p$ for points, which are just integer pairs.

**Definition 1.** Given a point $p_1 = (i_1, j_1)$, another point $p_2 = (i_2, j_2)$ is a successor of $p_1$ if (1) $i_2 = i_1 + 1$ and $j_2 = j_1$, or (2) $i_2 = i_1$ and $j_2 = j_1 + 1$, or (3) $i_2 = i_1 + 1$ and $j_2 = j_1 + 1$.

**Definition 2.** Let $P = \{p_1, \ldots, p_R\}$ be a sequence of points such that $p_r \in [m] \times [n]$ holds for each $1 \leq r \leq R$. We call $P$ a path of order $(m, n)$ if $p_{r+1}$ is a successor of $p_r$ for each $1 \leq r < R$ (and this $P$ is sometimes also called a “warping path” [21]). Also, we refer to a path of length 2 as a step that connects a point to one of its successors.

We use $\mathcal{P}(m, n)$ for the set of paths of order $(m, n)$. A path $P_1 \in \mathcal{P}(m, n)$ is a subpath of another path $P_2 \in \mathcal{P}(m, n)$ if $P_1$ is contained in $P_2$ (as a consecutive segment).

**Definition 3.** Let $P_1$ and $P_2$ be two non-empty paths. We use $P_1 \simeq P_2$ to mean that $P_1$ and $P_2$ begin at the same point and end at the same point.

**Definition 4.** Let $P_1$ and $P_2$ be two paths such that the first point of $P_2$, if it exists, is the successor of the last point of $P_1$, if it exists. We write $P_1 + P_2$ to mean the concatenation of $P_1$ and $P_2$ (as sequences of points) that forms a path containing both $P_1$ and $P_2$ as its subpaths. In the case where both $P_1$ and $P_2$ are non-empty, there is a step in $P_1 + P_2$ connecting $P_1$ and $P_2$ that consists of the last point in $P_1$ and the first point in $P_2$.

Also, we write $P_1 \oplus P_2$ to mean $P_1 + P_2'$ where the last point of $P_1$ is assumed to be the first point of $P_2$ and $P_2'$ is the tail of $P_2$, that is, $P_2'$ is obtained from removing the first point in $P_2$. In other words, $P_1 + P_2$ implies that $P_1$ and $P_2$ share no point while $P_1 \oplus P_2$ implies that $P_1$ and $P_2$ share one point, which is the last point of $P_1$ and the first point of $P_2$.

### Figure 1

An optimal full path of the order $(10, 6)$ whose cost equals 1.

|   |   |   |   |   |   |   |   |   |   |
|---|---|---|---|---|---|---|---|---|---|
| d | 3 | 3 | 3 | 2 | 2 | 2 | 0 | 0 | 0 |
| d | 3 | 3 | 3 | 2 | 2 | 2 | 0 | 0 | 0 |
| c | 2 | 2 | 2 | 1 | 1 | 1 | 1 | 1 | 1 |
| b | 1 | 1 | 1 | 0 | 0 | 0 | 0 | 2 | 2 |
| a | 0 | 0 | 0 | 1 | 1 | 1 | 1 | 3 | 3 |
| a | 0 | 0 | 0 | 1 | 1 | 1 | 1 | 3 | 3 |

### Table

| $(x)$ |   |   |   |   |   |   |   |   |   |
|---|---|---|---|---|---|---|---|---|---|
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |
|   |   |   |   |   |   |   |   |   |   |

### Equation

\[(y) \cdot \]
Definition 5. Let \( x = (a_1, \ldots, a_m) \) and \( y = (b_1, \ldots, b_n) \) be two strings. For each path \( P \in \mathcal{P}(m,n) \), there is a value \( \text{cost}_{x,y}(P) = \sum_{r=1}^{R} \delta(a_{i_r}, b_{j_r}) \), where \( P \) equals \((i_1, j_1), \ldots, (i_R, j_R)\). This value is often referred to as the cost of \( P \). We may write \( \text{cost}(P) \) for \( \text{cost}_{x,y}(P) \) if it is clear from the context what \( x \) and \( y \) should be.

We call each \( P \in \mathcal{P}(m,n) \) a full path if \((i_1, j_1) = (1,1)\) and \((i_R, j_R) = (m,n)\). The DTW distance between \( x \) and \( y \), denoted by \( \text{DTW}(x,y) \), is formally defined as the minimum of \( \text{cost}_{x,y}(P) \), where \( P \) ranges over the set of full paths of order \((m,n)\). Also, a full path \( P \) is referred to as an optimal full path if \( \text{cost}_{x,y}(P) = \text{DTW}(x,y) \). Given there are only finitely many paths of order \((m,n)\), there must exist one full path that is optimal. As an example, the shaded squares in Figure 1 illustrate the following full path of the order \((10,6)\):

\[
((1,1), (2,2), (3,2), (4,3), (5,3), (6,3), (7,4), (8,5), (9,6), (10,6))
\]

where the number in each square is the assumed distance between the two corresponding letters (computed here as the difference between their positions in the alphabet).

Definition 6. Let \( P = ((i_1, j_1), \ldots, (i_R, j_R)) \).

1. \( P \) is a \( v \)-path if all the \( i_r \) are the same for \( 1 \leq r \leq R \).
2. \( P \) is a \( h \)-path if all the \( j_r \) are the same for \( 1 \leq r \leq R \).
3. \( P \) is a \( d \)-path if \( i_{r+1} = i_r + 1 \) and \( j_{r+1} = j_r + 1 \) for \( 1 \leq r < R \).

Please recall that a step is a path of length 2. If a step is a \( h \)-path/\( v \)-path/\( d \)-path, respectively, then it is a \( h \)-step/\( v \)-step/\( d \)-step, respectively.

Definition 7. Let \( x = (a_1, \ldots, a_m) \) and \( y = (b_1, \ldots, b_n) \) be two strings. We use \( T_{\text{DTW}}(x,y) \) for the table \( T_{m,n} \times 1 \) such that the content of \( T_{m,n}[i,j] \) is \( \delta(a_i, b_j) \) for each \( i \in [m] \) and \( j \in [n] \).

We may use \( T_{\text{DTW}} \) for \( T_{\text{DTW}}(x,y) \) if \( x \) and \( y \) can be readily inferred from the context. If a table \( T_{\text{DTW}} \) can be readily inferred from the context, we often associate a point \((i,j)\) with the entry \( T_{\text{DTW}}[i,j] \) and think of a path \( P = ((i_1, j_1), \ldots, (i_R, j_R)) \) as the sequence of entries \( T_{\text{DTW}}[i_r, j_r] \) for \( 1 \leq r \leq R \). As an example, a full path of the order \((10,6)\) is given in Figure 1, where the path is indicated with the 10 shaded entries.

Suppose that the \( i \)-th run (\( j \)-th) in \( x \) (\( y \)) consists of the letters in \( x \) (\( y \)) from position \( i_1 \) (\( j_1 \)) to position \( i_2 \) (\( j_2 \)), inclusive. Then there is a corresponding block \( B_{i,j} \) consisting of all the entries \( T_{\text{DTW}}[u,v] \) for \( i_1 \leq u \leq i_2 \) and \( j_1 \leq v \leq j_2 \). Finally, we introduce notation for discussing specific blocks:

Definition 8. If there exists a block \( B \) to the right of \( B_{i,j} \) such that \( \delta(B) < \delta(B_{i,j}) \), we use \( \beta_0(B_{i,j}) \) for such a \( B \) that is the closest to \( B_{i,j} \). In other words, \( \beta_0(B_{i,j}) = B_{i',j'} \) for the least \( i' \) satisfying \( i < i' \) and \( \delta(B_{i',j'}) < \delta(B_{i,j}) \). Similarly, if there exists a block \( B \) above \( B_{i,j} \) such that \( \delta(B) < \delta(B_{i,j}) \), then \( \beta_v(B_{i,j}) \) is \( B_{i,j'} \) for the least \( j' \) satisfying \( j < j' \) and \( \delta(B_{i,j'}) < \delta(B_{i,j}) \).

### 3.1 Computing DTW Distance with Graphs

It is well known [40] that one can turn the problem of computing \( \text{DTW}(x,y) \) for two given strings \( x \) and \( y \) into a problem of finding the shortest distance between two given vertices in some graph, as follows.

Let \( x = (a_1, \ldots, a_m) \) and \( y = (b_1, \ldots, b_n) \). We can construct a directed graph \( G_0 = (V_0, E_0) \) such that

1. there is a vertex \( v_{i,j} \in V_0 \) for each pair \((i,j) \in [m] \times [n]\), and
2. there is a directed edge \((v_{i_1,j_1}, v_{i_2,j_2})\) of length \( \delta(a_{i_1}, b_{j_1}) \) connecting \( v_{i_1,j_1} \) to \( v_{i_2,j_2} \) whenever \((i_2,j_2)\) is a successor of \((i_1,j_1)\).
We use $G_{DTW}(x,y)$ for this graph $G_0$ and use $v$ and $e$ to range over $V_0$ and $E_0$, respectively. We may also refer to each vertex $v_{i,j} \in V_0$ simply as point $(i,j)$ if there is no risk of confusion. Clearly, $|V_0|$, the size of $V_0$, is $mn$, and $|E_0|$, the size of $E_0$, is bounded by $3mn$ (since each point can have at most 3 successors).

As every warping path is naturally mapped to a path in the graph $G_{DTW}(x,y)$ and vice versa, we can use $P$ to range over both warping paths in $T_{DTW}(x,y)$ and paths in $G_{DTW}(x,y)$ without risking confusion. Given a (non-empty) warping path $P$ in $T_{DTW}(x,y)$, we use $\text{len}(P)$ for the length of the corresponding path of $P$ in $G_{DTW}(x,y)$, which equals the cost of $P$ minus the cost associated with the last point in $P$. Therefore, finding the value of $DTW(x,y)$ is equivalent to finding the shortest distance from $v_{1,1}$ to $v_{m,n}$, which can be done by running some version of Dijkstra’s shortest distance algorithm. Alternatively, since $G_{DTW}(x,y)$ is acyclic, one can use dynamic programming to find the shortest distance, in which case the running time becomes $O(mn)$. This yields the classic dynamic-programming solution for computing $DTW(x,y)$ [40].

The basic strategy that we use in this paper to design a DTW approximation algorithm can be outlined as follows. Let $G_0 = \langle V_0, E_0 \rangle$ be the graph $G_{DTW}(x,y)$ given above. We try to construct a graph $G = \langle V, E \rangle$ such that $V \subseteq V_0$ holds and the length of each edge $e$ in $E$ that connects a vertex $v_1$ to another vertex $v_2$ equals the shortest distance from $v_1$ to $v_2$ as is defined in $G_0$. Let $\text{dist}_0$ and $\text{dist}$ be the shortest distance functions on the graphs $G_0$ and $G$, respectively. We attempt to prove that

$$\text{dist}_0(v_{1,1}, v_{m,n}) \leq \text{dist}(v_{1,1}, v_{m,n}) \leq \alpha \cdot \text{dist}_0(v_{1,1}, v_{m,n})$$

for some approximation ratio $\alpha > 1$ (e.g., $\alpha = 1 + \epsilon$ for $\epsilon > 0$). By running a shortest-path algorithm on $G$, we are able to compute $\text{dist}(v_{1,1}, v_{m,n})$ and thus obtain an $\alpha$-approximation algorithm for $DTW(x,y)$. As the time complexity of such an algorithm can be bounded by $O(|E|)$ plus the time needed for constructing $G$, the key to finding a fast algorithm is to try to minimize $|E|$, the size of $E$ (while ensuring that the construction of $G$ can be done in $O(|E|)$ time).

### 4 A $(1+\epsilon)$-Approximation Algorithm for DTW

In this section, we present and analyze a $(1+\epsilon)$-approximation algorithm for approximating the DTW distance between two run-length encoded strings in near-quadratic time.

Let $x = (a_1, \ldots, a_m)$ and $y = (b_1, \ldots, b_n)$ be two non-empty strings. Following Section 3.1, our approach will be to construct a graph $G = \langle V, E \rangle$ based on $G_{DTW}(x,y)$, reducing the problem of computing a $(1+\epsilon)$-approximation of $DTW(x,y)$ to finding the shortest distance between the vertex $v_{1,1}$ and the vertex $v_{m,n}$ in $G$.

We begin by describing the notions of h-to-v paths and v-to-h paths. The role that these will play in our algorithm is that we will show how to decompose any full path $P$ into a concatenation $P_1 + P_2 + \cdots$ of h-to-v and v-to-h paths.

**Definition 9.** Let $x$ and $y$ be two non-empty strings. A horizontal-to-vertical (h-to-v) path in $T_{DTW}(x,y)$ is a path that connects a point on the lower boundary of a block $B_{i,j}$ to another point on the right boundary of $B_{i,j}$. An h-to-v component in a full path is a maximal h-to-v path that is not contained in any longer h-to-v path in the same full path. A vertical-to-horizontal (v-to-h) path can be defined similarly.
Note that an h-to-v path matches characters from a single run in \( x \) to characters from (possibly multiple) runs in \( y \). We now make the (standard) observation that, when we are comparing a single run to characters to a multi-run string, DTW behaves in a very natural way:

**Observation 10.** Let \( x = (a_1, \ldots, a_m) \) and \( y = (b_1, \ldots, b_n) \) be two non-empty strings. Assume that \( x \) is a run of some letter \( a_0 \), that is, \( a_0 = a_i \) for \( 1 \leq i \leq m \).
1. If \( m \leq n \), then we have: \( DTW(x, y) = \sum_{j=1}^{n} \delta(a_0, b_j) \). This case corresponds to a path of the form \( D \oplus U \), where \( D \) consists of only \( d \)-steps and \( U \) only \( v \)-steps.
2. If \( m \geq n \), then we have: \( DTW(x, y) = \sum_{j=1}^{m} \delta(a_0, b_j) + (m - n) \cdot \delta(a_0, b_0) \), where \( b_0 \) is some \( b_j \) closest to \( a_0 \), that is, \( \delta(a_0, b_0) = \min(\delta(a_0, b_j)) \) for \( 1 \leq j \leq n \). This case corresponds to a path of the form \( D_1 \oplus H \oplus D_2 \), where \( D_1 \) and \( D_2 \) consist of only \( d \)-steps and \( H \) only \( h \)-steps. It should be further noted that, in this case, \( H \) can be assumed to travel along the lower boundary of some block, without loss of generality.

We say that a path connecting \( p \) and \( q \) is optimal if its cost is the least among all the paths connecting \( p \) and \( q \). By merging the two cases in Observation 10, we can assume that each optimal h-to-v path is of the form \( D_1 \oplus H \oplus D_2 \oplus U \), where any of the four sub-components can vanish. From now on, we can use a 5-tuple \((p_1, q_1, q_2, p_2, p_3)\) (which may also be written as \( p_1 q_1 q_2 p_2 p_3 \)) to refer to an h-to-v path, where \( p_1 q_1 \) is \( D_1 \), \( q_1 q_2 \) is \( H \), \( q_2 p_2 \) is \( D_2 \), and \( p_2 p_3 \) is \( U \). Similarly, each optimal v-to-h path is of the form \( D_1 \oplus U \oplus D_2 \oplus H \), and we use a corresponding 5-tuple representation to refer to a v-to-h path as well.

Next we argue that any full path \( P \) can be decomposed into h-to-v and v-to-h paths.

**Lemma 11.** Let \( x \) and \( y \) be two non-empty strings. Given a full path \( P_0 \) in \( T_{DTW}(x, y) \), we have \( P_0 = P_1 + \cdots + P_R \) where \( P_r \) is an h-to-v path for each odd \( 1 \leq r \leq R \) and \( P_r \) is a v-to-h path for each even \( 1 \leq r \leq R \).

**Proof.** The proof follows directly from the definitions of h-to-v paths and v-to-h paths. For brevity, we defer the full proof to the extended version of the paper [41].

Given two non-empty strings \( x = (a_1, \ldots, a_m) \) and \( y = (b_1, \ldots, b_n) \), we outline as follows a strategy for approximating \( DTW(x, y) \). Let \( P_0 \) be an optimal full path on \( T_{DTW}(x, y) \) such that \( cost(P_0) = DTW(x, y) \). By Lemma 11, we have \( P_0 = P_1 + \cdots + P_R \), where \( P_1 \) is an h-to-v path and \( P_1, \ldots, P_R \) are a sequence of alternating h-to-v paths and v-to-h paths. Let us choose an h-to-v path \( P_r \) for some \( 1 \leq r \leq R \). By Observation 10, we can assume that \( P_r \) is of the form of solid lines depicted in Figure 2.\(^5\)

\(^5\) The meaning of the dashed lines in the figure is to be explained later.
In more detail, the path $P_t$ moves diagonally from a point $p_1$ on the lower boundary of a block $B_1$ until it meets the lower boundary of another block; it moves horizontally along that lower boundary for some distance; it then moves diagonally to reach a point $p_2$ on the right boundary of another block $B_2$ (which is either $B_1$ or sits above $B_1$); and finally it moves vertically to reach a point $p_3$ on the right boundary of another block $B_3$ (which is either $B_2$ or sits above $B_2$). Note that the horizontal moves contained in $P_t$ must be inside a block where those moves cost the least.\footnote{If there are several blocks in which such horizontal moves can take place, we simply assume that the moves are inside the lowest of these blocks.}

Let $G_0 = \langle V_0, E_0 \rangle$ be the graph $G_{DTW}(x, y)$ described in Section 3.1 for computing $DTW(x, y)$. We may use a point (that is, an integer pair) to refer to the corresponding vertex in $V_0$. We may also use a path $P$ in $T_{DTW}(x, y)$ to denote its counterpart in $G_0$.

Given $\epsilon > 0$, we will construct a graph $G = \langle V, E \rangle$ such that:

- The point $(1, 1)$ is in $V$ and $V \subseteq V_0$ holds.
- Every point in $V$ is on a boundary. Each point in $V$ that is on either a right or upper boundary is connected by an edge to the next snap point on the same boundary (if there is one).
- If there is a step (either an h-step, v-step, or d-step) connecting two boundary points $p_1$ and $p_2$ in $E_0$, then there is also a step connecting $snap(p_1)$ and $snap(p_2)$ in $E$, where $snap(p_1)$ (resp. $snap(p_2)$) is the nearest point in $V$ above or to the right of $p_1$ (resp. $p_2$) on the same boundary as $p_1$.
- For each h-to-v (resp. v-to-h) path $P$ from $p_1$ to $p_2$ (depicted by some solid lines in Figure 2) in $G_0$ and any point $p_1'$ in $G$ to the right of $p_1$ (resp. above $p_1$) such that $p_1$ and $p_1'$ are on the same block boundary, there exists a path $P'$ (depicted by some dashed lines in Figure 2) in $G$ connecting $p_1'$ and the point $p_2' = snap(p_2)$ in $G$ such that $dist(P') \leq (1 + \epsilon)dist_0(P)$, where $dist$ and $dist_0$ are the shortest distance functions on the graphs $G$ and $G_0$, respectively.

We remark that our construction of $G$ will repeatedly make use of the following basic fact.

\textbf{Observation 12.} Let $\Delta(t) = [(1 + \epsilon)^t]$ for integers $t \geq 0$. For each integer $d \geq 1$, we have a $(1 + \epsilon)$-approximation of $d$ that is of the form $\Delta(t)$. In other words, $d \leq \Delta(t) \leq (1 + \epsilon) \cdot d$ holds for some $t$.

We now describe how to construct the graph $G$. We construct the set $V$ of vertices as follows:

1. Each vertex in $G_0$ corresponding to a corner point in $T_{DTW}(x, y)$ should be added into $V$. There are at most $4k\ell$ such vertices, where $k = \|x\|$ and $\ell = \|y\|$. 
2. Assume $(i, j)$ is the lower-left corner of block $B$.
   - If a point $(i + 1 + \Delta(t), j)$ is on the lower boundary of $B$ for some $t \geq 0$, then this point should be added into $V$. There are at most $\log_{1+\epsilon}(m)$ such points for the block $B$.
   - If a point $(i, j + 1 + \Delta(t))$ is on the left boundary of $B$ for some $t \geq 0$, then this point should be added into $V$. There are at most $\log_{1+\epsilon}(n)$ such points for the block $B$.
3. Assume $(i, j)$ is the upper-left corner of block $B$. If a point $(i + \Delta(t), j)$ is on the upper boundary of $B$ for some $t \geq 0$, then this point should be added into $V$. There are at most $\log_{1+\epsilon}(m)$ such points for the block $B$.
4. Assume $(i, j)$ is the lower-right corner of block $B$. If a point $(i, j + \Delta(t))$ is on the right boundary of $B$ for some $t \geq 0$, then this point should be added into $V$. There are at most $\log_{1+\epsilon}(n)$ such points for the block $B$. 

\textbf{Observation 12.} Let $\Delta(t) = [(1 + \epsilon)^t]$ for integers $t \geq 0$. For each integer $d \geq 1$, we have a $(1 + \epsilon)$-approximation of $d$ that is of the form $\Delta(t)$. In other words, $d \leq \Delta(t) \leq (1 + \epsilon) \cdot d$ holds for some $t$.
The points in $V$ are referred to as **snap points**. Given a snap point $p'$ on the upper or right boundary of some block, if $p'q'$ is a d-step for some point $q'$, then $q'$ is also a snap point. This can be readily verified by inspecting the construction of $V$. Also, for each block $B$, there are at most $O(\log(m + n)/\epsilon)$ points added to $V$. Therefore, $|V|$, the size of $V$, is $O(k\ell \cdot \log(m + n)/\epsilon)$ or simply $\tilde{O}(k\ell/\epsilon)$.

**Definition 13.** Given a point $p \in V_0$ on a horizontal boundary of a block $B$, we use $\text{snap}_h(p)$ for the point $p' \in V$ such that $p'$ is $p$ if $p \in V$ or $p'$ is the closest point to the right of $p$ that is on the same boundary of $B$. The existence of such a point is guaranteed as all of the corner points are included in $V$. Similarly, $\text{snap}_v(p)$ can be defined for each point $p$ on a vertical boundary of a block.

We can use $\text{snap}(p)$ for either $\text{snap}_h(p)$ or $\text{snap}_v(p)$ without confusion: If both $\text{snap}_h(p)$ and $\text{snap}_v(p)$ are defined for $p$, then $p$ must be a corner of some block $B$, implying $p = \text{snap}_h(p) = \text{snap}_v(p)$ since $p \in V$ holds. We argue as follows that finding $\text{snap}(p)$ for each given $p$ can be done in $\tilde{O}(1)$ time.

**Definition 14.** Let $x = (a_1, \ldots, a_m)$ be a string and $\hat{x} = (a'_1 \cdot m_1, \ldots, a'_k \cdot m_k)$ be its RLE representation. Let $M_r$ be $m_1 + \cdots + m_r$ for each $0 \leq r < k$. For each $1 \leq i \leq m$, we use $\hat{i}_x$ for the pair $(i_0, i_1)$ such that $i = M_{i_0} + i_1$ for $1 \leq i_1 \leq m_{i_0 + 1}$.

We may use $\hat{i}$ for $\hat{i}_x$ if $x$ can be readily inferred from the context.

It is worth taking a moment to verify that we can compute $\hat{j}_x$ efficiently. Assume that an array storing $M_r$ for $0 \leq r < k$ is already built (in $O(k)$ time). Given $i \in [m]$, we can perform binary search on the array to find $i_0$ in $O(\log(k))$ time such that $M_{i_0} < i \leq M_{i_0 + 1}$; we can then compute $\hat{i}_x$ as $(i_0, i - M_{i_0})$.

It is also worth verifying that we can compute $\text{snap}(p)$ in $\tilde{O}(1)$ time. Given a point $p = (i, j)$ on a boundary of some block $B$ in $T_{DTW}(x, y)$, we can compute $\hat{i}_x = (i_0, i_1)$ in $O(\log(k))$ time. Similarly, we can compute $\hat{j}_y = (j_0, j_1)$ in $O(\log(l))$ time. We can locate the block $B$ as $B_{i_0 + 1, i_1 + 1}$, and then find $\text{snap}(p)$ in $O(1)$ time (assuming $\log(l)$ and $\log(l)$ can be computed in $O(1)$ time). Therefore, given $p$, we can compute $\text{snap}(p)$ in $\tilde{O}(1)$ time.

Having established that we can compute $\hat{i}$ and $\text{snap}(p)$ efficiently, we are nearly ready to describe the construction of the edges. Our final task before doing so is to establish a bit more notation for how to talk about blocks.

Please recall that $\beta_h(B_{i,j})$ (resp. $\beta_v(B_{i,j})$) refers to the closest block $B_{i',j}$ (resp. $B_{i,j'}$) such that $\delta(B_{i',j}) < \delta(B_{i,j'})$ (resp. $\delta(B_{i,j'}) < \delta(B_{i,j})$) holds. If there is no such a block, $\beta_h(B_{i,j})$ (resp. $\beta_v(B_{i,j})$) is undefined.

**Definition 15.** We refer to $B_{i_1, j_1}, \ldots, B_{i_s, j_s}$ as a $\beta_h$-sequence if $B_{i_{s+1}, j} = \beta_h(B_{i_s, j})$ for $1 \leq s < S$. Let $\beta_h(x, y)$ be the length of a longest $\beta_h$-sequence. Clearly, we have $\beta_h^*(x, y) \leq ||x||$.

Similarly, we refer to $B_{i_1, j_1}, \ldots, B_{i_s, j_s}$ as a $\beta_v$-sequence if $B_{i, j_{s+1}} = \beta_v(B_{i, j_s})$ for $1 \leq s < S$. Let $\beta_v^*(x, y)$ be the length of a longest $\beta_v$-sequence. Clearly, we have $\beta_v^*(x, y) \leq ||y||$.

Let $\beta^*(x, y) = \max(\beta_h^*(x, y), \beta_v^*(x, y))$.

Observe that if the underlying distance function $\delta$ on letters is from Hamming space, then $\beta^*(x, y) \leq 2$ for any $x$ and $y$. Slightly more generally, if $\delta$ is bounded by a constant, then $\beta^*(x, y)$ is bounded by the same constant plus one. Later in the section, in the proof of Theorem 21, we will also see an important (and much more general) case where $\beta^*(x, y)$ is guaranteed to be $\tilde{O}(1)$.

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7 We slightly abuse the $\tilde{O}$ notation here as the parameters $m$ and $n$ for the implicit log-terms are not explicitly mentioned.
We are now ready to explain the construction of the set $E$ of edges for connecting vertices in $V$. The basic idea is to construct $E$ in such a way that, for any constructed path $p_1'q_2'q_3'q_4'$ (as is depicted in Figure 2), there should be a path in $G$ going from $p_1'$ to $p_2'$ whose cost is at most the cost of the path in $G_0$ — this allows for the graph $G$ to capture all such paths, and ultimately allows for $G$ to be used in our approximation algorithm. Formally, the construction of $E$ can be performed with the following steps:

1. Note that the corner points in $V_0$ are all in $V$. The edges connecting these corner points in $E_0$ should be added into $E$.

2. Given a point $p_1' \in V$ on the upper boundary of a block $B$, if there is a d-step from $p_1'$ to $p_2'$ (on the lower boundary of the block above $B$), then $p_2'$ is in $V$ and an edge from $p_1'$ to $p_2'$ should be added into $E$ whose length equals $\delta(B)$.

3. Given a point $p_1' \in V$ on the right boundary of a block $B$, if there is a d-step from $p_1'$ to $p_2'$ (on the left boundary of the block to the right of $B$), then $p_2'$ is in $V$ and an edge from $p_1'$ to $p_2'$ should be added into $E$ whose length equals $\delta(B)$.

4. If two points $p_1' = (i_1,j_1)$ and $p_2' = (i_2,j_2)$ in $V$ are on the same horizontal or vertical boundary of a block $B$ such that $p_2'$ is the closest point above or to the right of $p_1'$, then an edge from $p_1'$ to $p_2'$ should be added into $E$ whose length equals $(i_2 - i_1) \cdot \delta(B)$ (horizontal) or $(j_2 - j_1) \cdot \delta(B)$ (vertical).

5. Let $p_1'$ be a point in $V$ on the lower boundary of a block $B_1$. This step adds into $E$ edges between $p_1'$ and certain chosen snap points $q_1'$ such that there are h-to-v paths connecting $p_1'$ and $q_1'$.

6. Let us use $B_1^1, \ldots, B_1^S$ for the sequence where $B_1 = B_1^1$ and $B_1^{s+1} = \beta_s(B_1^s)$ for $1 \leq s < S$ and $\beta_s(B_1^S)$ is undefined. Clearly, $S$ is bounded by $\beta_1^S(x, y)$ (according to the definition of $\beta_1^S(x, y)$). Let $B_1^1$ range over $B_1^1, \ldots, B_1^S$.

7. Let $q_1' = (i_1', j_1')$ be the point on the lower boundary of $B_1^1$ such that the path connecting $p_1'$ and $q_1'$ consists of only d-steps. Let $\text{snaps}_d(q_1')$ be the set consisting of the point $q_1'$, the points on the lower boundary of $B_1^1$ of the form $(i_1' + \Delta(t), j_1')$ for some $t \geq 0$, and the lower-right corner point of $B_1^1$. For each $q_2'$ ranging over the set $\text{snaps}_d(q_1')$, there exists at most one point $q_3'$ on the right boundary of some $B_2$ (which is either $B_1^1$ or sits above $B_1^1$) such that the path connecting $q_2'$ and $q_3'$ consists of only d-steps. As this $q_3'$ may not be in $V$, we choose $q_3'$ to be $\text{snaps}_d(q_3')$, which is in $V$ by definition. Note that the path $p_1'q_1'q_2'q_3'q_4'$ is an h-to-v path in $T_{DTW}(x, y)$. We add into $E$ an edge between $p_1'$ and $q_4'$ for each $q_4'$. The length of each added edge connecting $p_1'$ and $q_4'$ is the shortest distance between $p_1'$ and $q_4'$, which, by Lemma 17, can be computed in $O(1)$ time.

8. There is one $q_4'$ for each $B_1^t$, and there are at most $\log_{1+\epsilon}(m)$ many of $q_2'$ for each $q_1'$, and there is at most one $q_3'$ for each $q_2'$ and one $q_4'$ for each $q_3'$. Therefore, for each $p_1'$, there are at most $\beta_1^S(x, y) \cdot \log_{1+\epsilon}(m)$ edges added into $E$.

9. Let $p_1'$ be a point in $V$ on the left boundary of a block $B_1$. This step adds into $E$ edges between $p_1'$ and certain chosen snap points $q_4'$ such that there are v-to-h paths connecting $p_1'$ and $q_4'$. We omit the details that are parallel to those in the previous step. There are at most $\beta_1^S(x, y) \cdot \log_{1+\epsilon}(n)$ edges added into $E$ for each $p_1'$.

Let us take a moment to discuss how to efficiently compute the lengths of the edges added to $E$ during the construction of $G = (V, E)$. That is, how to construct and determine the cost of each dotted path $p_1'q_2'q_3'q_4'$ depicted in Figure 2. (Note that $q_4' = \text{snaps}_v(q_4')$ is not shown in the figure.)

**Lemma 16.** Let $x$ and $y$ be two non-empty strings. For $k = ||x||$ and $\ell = ||y||$,

1. we can compute $\beta_d(B)$ for all the blocks $B$ in $T_{DTW}(x, y)$ in $O(k \ell \cdot \log(k))$ time, and

2. we can compute $\beta_v(B)$ for all the blocks $B$ in $T_{DTW}(x, y)$ in $O(k \ell \cdot \log(\ell))$ time.
Proof. We defer the full proof to the extended version of the paper [41].

Lemma 17. For each h-to-v path (p'_1, q'_1, q'_2, q'_3, q'_4), its length can be computed in $O(1)$ time if the five snap points p'_1, q'_1, q'_2, q'_3, and q'_4 are given.

Proof. We defer the full proof to the extended version of the paper [41].

For brevity, we omit the obvious lemma parallel to Lemma 17 that is instead on computing the lengths of v-to-h paths in $O(1)$ time.

We are now in a position to state and prove the main theorems of the paper. As noted earlier, the basic idea behind our $(1 + \epsilon)$-approximation algorithms is to compute a path-distance through the graph $G = (V, E)$, and show that this distance closely approximates $DTW(x, y)$.

We begin by stating a theorem that parameterizes its running time by $\beta^*(x, y) – we will then apply this result to obtain fast running times in the cases where the distance function $\delta$ outputs either $O(\log n)$-bit integer values (Theorem 21) or $\{0, 1\}$-values (Theorem 22).

Theorem 18. Let $x = (a_1, \ldots, a_m)$ and $y = (b_1, \ldots, b_n)$ be two non-empty strings, and let $\hat{x}$ and $\hat{y}$ denote the run-length encoded versions of the two strings. There exists a $(1 + \epsilon)$-approximation algorithm (ApproxDTW) for each $\epsilon > 0$ that takes $\hat{x}$ and $\hat{y}$ as its input and returns a value $\hat{DTW}(x, y)$ satisfying $DTW(x, y) \leq \hat{DTW}(x, y) \leq (1 + \epsilon) \cdot DTW(x, y)$. Moreover, the worst-case time complexity of this algorithm is $O(k \cdot \beta^*(x, y)/\epsilon^2)$ for $k = \|x\|$ and $\ell = \|y\|$, where $\beta^*(x, y)$ is defined in Definition 15.

Proof. The analysis of the approximation ratio follows as in Section 2.

4.1 Time-Bound for Polynomially-Bounded Letter Distances

In this section, we present a variant of the algorithm ApproxDTW for approximating $DTW(x, y)$ under the general condition that the distances between letters are integer values bounded by some polynomial of the lengths of $x$ and $y$. The time complexity of this variant, which takes $\hat{x}$ and $\hat{y}$ as its input to compute $DTW(x, y)$, is $O(k \epsilon/\ell^3)$ for $k = \|x\|$ and $\ell = \|y\|$.

Definition 19. Let $\delta$ be a distance function on letters such that $\delta(a, b) \geq 1$ if $\delta(a, b) \neq 0$. Given $\epsilon_1 > 0$, we use $\delta_{\epsilon_1}$ for the distance function such that $\delta_{\epsilon_1}(a, b) = 0$ if $\delta(a, b) = 0$, or $\delta_{\epsilon_1}(a, b) = \text{cpow}(1 + \epsilon_1, \delta(a, b))$ if $\delta(a, b) \geq 1$, where $\text{cpow}(1 + \epsilon_1, \alpha)$ equals $(1 + \epsilon_1)^{\lfloor \alpha \rfloor}$ for the least integer $t$ such that $\alpha \leq (1 + \epsilon_1)^t$ holds.

Please note that $\delta(a, b) \leq \delta_{\epsilon_1}(a, b) \leq (1 + \epsilon_1) \cdot \delta(a, b)$ holds for any letters $a$ and $b$.

Lemma 20. Let $DTW(\delta)$ be the DTW distance function where the underlying distance function for letters is $\delta$. Given $\epsilon_1 > 0$, we have the following inequality for each pair of strings $x$ and $y$:

$$DTW(\delta_{\epsilon_1})(x, y) \leq (1 + \epsilon_1) \cdot DTW(\delta)(x, y)$$

Proof. Let $P$ be an optimal full path such that its length based on $\delta$ equals $DTW(\delta)(x, y)$. We know that the length of $P$ based on $\delta_{\epsilon_1}$ is bounded by $(1 + \epsilon_1) \cdot DTW(\delta)(x, y)$ since $\delta_{\epsilon_1}(a, b) \leq (1 + \epsilon_1) \cdot \delta(a, b)$ holds for any letters $a$ and $b$. As $DTW(\delta_{\epsilon_1})(x, y)$ is bounded by the length of $P$ based on $\delta_{\epsilon_1}$, we have the claimed inequality.
approximating DTW distance between RLE strings

Let $\epsilon_1 > 0$ and $\epsilon_2 > 0$. By Lemma 20, every $(1+\epsilon_2)$-approximation algorithm for DTW based on $\delta_{\epsilon_1}$ is a $(1+\epsilon_1) \cdot (1+\epsilon_2)$-approximation algorithm for DTW based on $\delta$. For each $\epsilon > 0$, if we choose, for example, $\epsilon_1 = \epsilon/2 - \epsilon^2/2$ and $\epsilon_2 = \epsilon/2$, then we have $(1+\epsilon_1) \cdot (1+\epsilon_2) < 1 + \epsilon$, implying that every $(1+\epsilon_2)$-approximation algorithm for DTW based on $\delta_{\epsilon_1}$ is a $(1+\epsilon)$-approximation algorithm for DTW based on $\delta$.

\[ \text{Theorem 21.} \] Let $x = (a_1, \ldots, a_m)$ and $y = (b_1, \ldots, b_n)$ be two non-empty strings.

Assume that the underlying distance function $\delta$ satisfies (1) $\delta(a, b) \geq 1$ if $\delta(a, b) \neq 0$ and (2) $\delta(a, b)$ is poly$(m+n)$ for any letters $a$ in $x$ and $b$ in $y$. There exists a $(1+\epsilon)$-approximation algorithm for each $\epsilon > 0$ that takes $\hat{x}$ and $\hat{y}$ as its input and returns a value $w$ satisfying $DTW(x, y) \leq w \leq (1 + \epsilon) \cdot DTW(x, y)$. And the worst-case time complexity of this algorithm is $\tilde{O}(k^2/\epsilon^3)$ for $k = \|x\|$ and $\ell = \|y\|$.

\[ \text{Proof.} \] Let $\epsilon_1 = \epsilon/2 - \epsilon^2/2$ and $\epsilon_2 = \epsilon/2$. By Theorem 18, ApproxDTW (as is presented in the proof of Theorem 18) takes $\hat{x}$ and $\hat{y}$ as input and returns a $(1 + \epsilon_2)$-approximation of $DTW(\delta_{\epsilon_1})(x, y)$. And the time complexity of the algorithm is $\tilde{O}(k^2/\epsilon_2^3)$.

Note that there are only $O(\log_{1+\epsilon_1}(m+n))$-many distinct values of $\delta_{\epsilon_1}(a, b)$ for $a$ and $b$ ranging over letters in $x$ and $y$, respectively. Hence, for the underlying distance function $\delta_{\epsilon_1}$ on letters, $\beta_{\delta_{\epsilon_1}}^*(x, y)$ is $O(\log_{1+\epsilon_1}(m+n))$ and $\beta_{\delta_{\epsilon_1}}^*(x, y)$ is also $O(\log_{1+\epsilon_1}(m+n))$, which implies that $\beta^*(x, y)$ is $O(\log_{1+\epsilon_1}(m+n))$ or simply $\tilde{O}(1/\epsilon_1)$. Therefore, we can use ApproxDTW to compute a $(1 + \epsilon_2)$-approximation of $DTW(\delta_{\epsilon_1})(x, y)$ in $\tilde{O}(k^2/\epsilon_1\epsilon_2^3)$ time. Since any $(1 + \epsilon_2)$-approximation of $DTW(\delta_{\epsilon_1})(x, y)$ is a $(1+\epsilon)$-approximation of $DTW(\delta)(x, y)$, we are done.

4.2 Time-Bound for Constant-Bounded Letter Distances

Assume that there exists a constant $N$ such that $\delta(a, b)$ is an integer less than $N$ for each pair $a$ and $b$ in $\Sigma$. For instance, $(\Sigma, \delta)$ satisfies this condition if it is Hamming space (for which $N$ can be set to 2).

\[ \text{Theorem 22.} \] Assume that $\delta(a, b)$ are $O(1)$ for $a, b \in \Sigma$. Then ApproxDTW, the $(1+\epsilon)$-approximation algorithm for DTW given in the proof of Theorem 18, runs in $O(k^2/\epsilon^2)$ time for $k = \|x\|$ and $\ell = \|y\|$, where $\hat{x}$ and $\hat{y}$ are the input of the algorithm.

\[ \text{Proof.} \] This theorem follows from Theorem 18 immediately since $\beta^*(x, y)$ is $O(1)$.

5 Conclusion

We have presented in this paper an algorithm for approximating the DTW distance between two RLE strings. Trading accuracy for efficiency, this algorithm is of (near) quadratic-time complexity and thus, as can be expected, asymptotically faster than the exact DTW algorithm of cubic-time complexity [21], which is currently considered the state-of-art of its kind.

It will be interesting to further investigate whether there exist asymptotically faster approximation algorithms for DTW than the one presented in this paper. In particular, it seems both interesting and challenging to answer the open question as to whether there exists a (near) quadratic-time algorithm for computing the (exact) DTW distance between two RLE strings.
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