

Algorithms for the Constrained Longest Common Subsequence Problems

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Abstract. Given strings S_1, S_2 , and P , the constrained longest common subsequence problem for S_1 and S_2 with respect to P is to find a longest common subsequence lcs of S_1 and S_2 such that P is a subsequence of this lcs . We present an algorithm which improves the time complexity of the problem from the previously known $O(rn^2m^2)$ to $O(rnm)$ where r, n , and m are the lengths of P, S_1 , and S_2 , respectively. As a generalization of this, we extend the definition of the problem so that the lcs sought contains a subsequence whose edit distance from P is less than a given parameter d . For the latter problem, we propose an algorithm whose time complexity is $O(drn)$.

Keywords: Longest common subsequence, constrained subsequence, edit distance, dynamic programming.

1 Introduction

A subsequence of a string S is obtained by deleting zero or more symbols of S . The *longest common subsequence* (lcs) problem for two strings is to find a common subsequence in both strings having maximum possible length. The lcs problem has many applications, and it has been studied extensively, see for example [1, 4, 2, 3, 5, 7]. The problem has a simple dynamic programming formulation. To compute an lcs between two strings of lengths n , and m , we use the *edit graph*. The edit graph is a directed acyclic graph having $(n + 1)(m + 1)$ lattice points (i, j) for $0 \leq i \leq n$, and $0 \leq j \leq m$ as vertices. Vertex $(0, 0)$ appears at the top-left corner, and the vertex (n, m) is at the bottom-right corner of this rectangular grid. To vertex (i, j) there are incoming arcs from its neighbors at $(i - 1, j)$, $(i, j - 1)$, and $(i - 1, j - 1)$ which represent, respectively, insert, delete, and either substitute or match operations. The lcs calculation counts the number of matches on the paths from vertex $(0, 0)$ to (n, m) , and the problem aims to maximize this number. The time complexity lower bound

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for the problem is $\Omega(n^2)$ for $n \geq m$ if the elementary operations are “equal/unequal”, and the alphabet size is unrestricted [1]. If the alphabet is fixed the best known time complexity is $O(n^2/\log n)$ when $n = m$ [5]. A survey of practical *lcs* algorithms can be found in [2].

Given strings S_1 , S_2 , and P , the constrained longest common subsequence problem [6] for S_1 and S_2 with respect to P is to find a longest common subsequence *lcs* of S_1 and S_2 such that P is a subsequence of this *lcs*. For example, for $S_1 = \text{bbaba}$, and $S_2 = \text{abbaa}$, bbaa is an (unrestricted) *lcs* for S_1 and S_2 , and aba is an *lcs* for S_1 and S_2 with respect to $P = \text{ab}$, as shown in Figure 1.

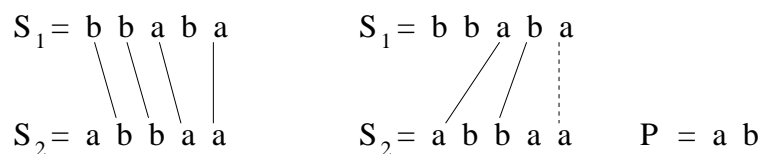


Figure 1: For $S_1 = \text{bbaba}$, and $S_2 = \text{abbaa}$, the length of an *lcs* is 4 (left). When constrained to contain $P = \text{ab}$ as a subsequence, the length of an *lcs* drops to 3 (right).

The problem is motivated by practical applications: For example in the computation of the homology of two biological sequences it is important to take into account a common specific or putative structure [6].

Let n, m, r denote the lengths of the strings S_1, S_2 , and P , respectively. Tsai [6] gave a dynamic programming formulation for the constrained longest common subsequence problem and a resulting algorithm whose time complexity is $O(rn^2m^2)$. In this paper we present a different dynamic programming formulation with which we improve the time complexity of the problem down to $O(rnm)$. We achieve improved results by changing the order of the dimensions in the formulation. We also extend the definition of the problem so that the *lcs* sought is forced to contain a subsequence whose *edit distance* from P is less than a given positive integer parameter d . For this latter problem we propose an algorithm whose time complexity is $O(drn^2m)$. Taking $d = 1$ specializes to the original constrained *lcs* problem as this choice of d forces the subsequence to contain P itself. We describe these results in section 2.

2 Algorithms

Let $|S_1| = n$, $|S_2| = m$ with $n \geq m$, and $|P| = r$. Let $S[i]$ denote the i th symbol of string S . Let $S[i..j] = S[i]S[i+1] \cdots S[j]$ be the substring of consecutive letters in S from position i to position j inclusive for $i \leq j$, and the empty string otherwise.

Denote by $L_{i,j,k}$ the length of an *lcs* for $S_1[1..i]$ and $S_2[1..j]$ with respect to $P[1..k]$. This simply means that the common subsequence is constrained to contain P as a subsequence in turn. We calculate the values $L_{i,j,k}$ by a dynamic programming formulation. Then $L_{n,m,r}$ is the length of an *lcs* of S_1 and S_2 containing P as a subsequence.

Theorem 1 For all i, j, k , $1 \leq i \leq n$, $1 \leq j \leq m$, $0 \leq k \leq r$, $L_{i,j,k}$ satisfies

$$L_{i,j,k} = \max\{L'_{i,j,k}, L_{i,j-1,k}, L_{i-1,j,k}\} \tag{1}$$

where

$$L'_{i,j,k} = \max\{L''_{i,j,k}, L'''_{i,j,k}\} \quad (2)$$

and

$$L''_{i,j,k} = \begin{cases} 1 + L_{i-1,j-1,k-1} & \text{if } (k = 1 \text{ or } (k > 1 \text{ and } L_{i-1,j-1,k-1} > 0)) \\ & \text{and } S_1[i] = S_2[j] = P[k] \\ 0 & \text{otherwise} \end{cases}$$

$$L'''_{i,j,k} = \begin{cases} 1 + L_{i-1,j-1,k} & \text{if } (k = 0 \text{ or } L_{i-1,j-1,k} > 0) \text{ and } S_1[i] = S_2[j] \\ 0 & \text{otherwise} \end{cases}$$

with boundary conditions $L_{i,0,k} = 0$, $L_{0,j,k} = 0$, for all i, j, k , $0 \leq i \leq n$, $0 \leq j \leq m$, $0 \leq k \leq r$.

Proof We prove the correctness of our formulation by induction on k for all i, j .

We will consider all possible ways of obtaining an *lcs* with respect to $P[1..k]$ at any node i, j . Essentially there are three cases to consider:

1. An *lcs* ending at the node $(i, j - 1)$ is extended with the horizontal arc $((i, j - 1), (i, j))$ ending at node (i, j) ,
2. An *lcs* ending at $(i - 1, j)$ is extended with the vertical arc $((i - 1, j), (i, j))$ ending at node (i, j) ,
3. An *lcs* ending at node $(i - 1, j - 1)$ is extended with the diagonal arc $((i - 1, j - 1), (i, j))$ ending at node (i, j) . In this case we distinguish between subcases depending on whether the diagonal arc is a matching for the given strings along with the pattern, or is a matching for the given strings only at the current indices.

The possible *lcs* extensions referred to in items 1 and 2 above are accounted for by $L_{i,j-1,k}$ and $L_{i-1,j,k}$ respectively in the statement of the theorem. The quantities $L''_{i,j,k}$ and $L'''_{i,j,k}$ in the statement of the theorem keep track of the two further possibilities described in item 3.

In the base case: when $k = 0$ (i.e. when P is the empty string) $L''_{i,j,k}$ is identically 0. Therefore $L'_{i,j,k} = L'''_{i,j,k}$ in (2). Since $k = 0$, the conjunction in the definition of $L'''_{i,j,k}$ is always satisfied. We see that putting $L_{i,j} = L_{i,j,0}$, (1) becomes

$$L_{i,j} = \max\{L'_{i,j}, L_{i,j-1}, L_{i-1,j}\}$$

where

$$L'_{i,j} = \begin{cases} 1 + L_{i-1,j-1} & \text{if } S_1[i] = S_2[j] \\ 0 & \text{otherwise} \end{cases}$$

which is the classical dynamic programming formulation for the ordinary *lcs* between S_1 and S_2 [7].

Assume that for $k - 1$ ($k \geq 1$), $L_{i,j,k-1}$ computed by (1) is the length of an *lcs* for $S_1[1..i]$ and $S_2[1..j]$ with respect to $P[1..k - 1]$ for all i, j and consider the calculation of $L_{i,j,k}$ when $k > 1$.

We define a *path* at node (i, j) as a simple path in the edit graph which includes at least one matching arc, starts at node $(0, 0)$, and ends at node (i, j) . A path with

respect to $P[1..k]$ includes matching diagonal arcs ending at a sequence of $k \geq 1$ distinct nodes $(a_1, b_1), (a_2, b_2), \dots, (a_k, b_k)$ such that for all ℓ , $1 \leq \ell \leq k$, $S_1[a_\ell] = S_2[b_\ell] = P[\ell]$. We define $\#match$ on a path as the number of matches between the symbols of S_1 , and S_2 , not necessarily involving symbols in P . An *lcs path* with respect to $P[1..k]$ ending at node (i, j) is a path with respect to $P[1..k]$ ending at node (i, j) with maximum $\#match$. Thus $L_{i,j,k}$ is $\#match$ on an *lcs path* at node (i, j) with respect to $P[1..k]$. Evidently $\#match = \#match(i, j, k)$ is a function of the indices i, j, k . We will omit these parameters when they are clear from the context.

We can extend any *lcs path* with respect to $P[1..k]$ ending at node $(i, j - 1)$ with the horizontal arc $((i, j - 1), (i, j))$ to obtain a path with respect to $P[1..k]$ ending at node (i, j) . Such an extension does not change $\#match$ on the path, and $L_{i,j,k} \geq L_{i,j-1,k}$.

Similarly we can extend any *lcs path* with respect to $P[1..k]$ ending at node $(i - 1, j)$ with the vertical arc $((i - 1, j), (i, j))$ to obtain a path with respect to $P[1..k]$ ending at node (i, j) . This extension does not change $\#match$ on the path either, and $L_{i,j,k} \geq L_{i-1,j,k}$. Therefore, $L_{i,j,k} \geq \max\{L_{i,j-1,k}, L_{i-1,j,k}\}$.

By using a matching arc $((i - 1, j - 1), (i, j))$, we can obtain paths with respect to $P[1..k]$ at node (i, j) by extending *lcs paths* with either respect to $P[1..k - 1]$, or with respect to $P[1..k]$ ending at node $(i - 1, j - 1)$. These two possibilities are accounted for by $L''_{i,j,k}$ and $L'''_{i,j,k}$ in the dynamic programming formulation, respectively.

First consider *lcs paths* with respect to $P[1..k - 1]$ ending at node $(i - 1, j - 1)$. We will show that $L''_{i,j,k}$ stores the maximum $\#match$ on paths obtained at node (i, j) by extending these paths.

If $S_1[i] = S_2[j] = P[k]$ then: If $k = 1$ then this is the first time the letter $P[1]$ appears as a matching arc on a path ending at node (i, j) since we are considering *lcs paths* with respect to $P[1..k - 1]$ ending at node $(i - 1, j - 1)$ and $S_1[i] = S_2[j] = P[1]$. Therefore, the *lcs length* relative to $P[1]$ at (i, j) is $L''_{i,j,1} = 1 + L_{i-1,j-1,0}$, which is one more than the length of an ordinary *lcs* between $S_1[1..i - 1]$ and $S_2[1..j - 1]$. If $k > 1$ and if there is an *lcs path* with respect to $P[1..k - 1]$ ending at node $(i - 1, j - 1)$ (i.e. if $L_{i-1,j-1,k-1} > 0$) then we can extend this path with a new match, and $\#match$ in the resulting path ending at node (i, j) becomes $L''_{i,j,k} = 1 + L_{i-1,j-1,k-1}$.

Next we consider *lcs paths* with respect to $P[1..k]$ ending at node $(i - 1, j - 1)$. We will show that $L'''_{i,j,k}$ stores the maximum $\#match$ on paths obtained at node (i, j) by extending these paths.

If $S_1[i] = S_2[j]$ then: Since the $k = 0$ case is considered earlier in the base case of the induction, we only consider the case when $k > 1$. If there is an *lcs path* with respect to $P[1..k]$ ending at node $(i - 1, j - 1)$ (i.e. if $L_{i-1,j-1,k} > 0$) then we can extend this path by adding a new match (which does not involve P), and $\#match$ in the resulting path relative to $P[1..k]$ ending at node (i, j) becomes $L'''_{i,j,k} = 1 + L_{i-1,j-1,k}$.

After setting $L'_{i,j,k} = \max\{L''_{i,j,k}, L'''_{i,j,k}\}$, the quantity $L'_{i,j,k}$ is equal to the maximum $\#match$ on paths with respect to $P[1..k]$ ending at node (i, j) ending with the arc $((i - 1, j - 1), (i, j))$. If there is no such path then $L'_{i,j,k} = 0$. Therefore $L_{i,j,k} \geq \max\{L'_{i,j,k}, L_{i,j-1,k}, L_{i-1,j,k}\}$.

From all possible *lcs paths* ending at neighboring nodes of (i, j) we can find their extensions ending at node (i, j) , and we can obtain an *lcs path* ending at node (i, j) with respect to $P[1..k]$ for all k . We calculate, and store in $L_{i,j,k}$ such *lcs lengths*. Now consider the structure of an *lcs path* with respect to $P[1..k]$ ending at node (i, j) . As

	b	b	a	b	a
a	0	0	1	1	1
b	1	1	1	2	2
b	1	2	2	2	2
a	1	2	3	3	3
a	1	2	3	3	4

$k = 0$

	b	b	a	b	a
a	0	0	1	1	1
b	0	0	1	2	2
b	0	0	1	2	2
a	0	0	3	3	3
a	0	0	3	3	4

$k = 1$

	b	b	a	b	a
a	0	0	0	0	0
b	0	0	0	2	2
b	0	0	0	2	2
a	0	0	0	2	3
a	0	0	0	2	3

$k = 2$

Figure 2: For $S_1 = \text{abbaa}$, $S_2 = \text{bbaba}$, and $P = \text{ab}$, the tables of values $L_{i,j,k}$ = the length of an *lcs* for $S_1[1..i]$ and $S_2[1..j]$ with respect to $P[1..k]$.

typical in dynamic programming formulations, we consider the possible cases of the last arc on such a path to obtain $L_{i,j,k} \leq \max\{L'_{i,j,k}, L_{i,j-1,k}, L_{i-1,j,k}\}$ which proves the theorem. •

Example: Figure 2 shows the contents of the dynamic programming tables for $S_1 = \text{bbaba}$, and $S_2 = \text{abbaa}$, and $P = \text{ab}$ for $k = 0, 1, 2$. For $k = 0$, the calculated values are simply the ordinary dynamic programming *lcs* table for S_1 and S_2 .

All $L_{i,j,k}$ can be computed in $O(rnm)$ time, using $O(rm)$ space using the formulation in Theorem 1 by noting that we only need rows $i-1$, and i during the calculations at row i . If actual *lcs* is desired then we can carry the *lcs* information for each k along with the calculations. This requires $O(rnm)$ space. By keeping track, on *lcs* for each k , of only the match points (i', j') of $P[u]$ for all u , $1 \leq u \leq r$, the space complexity can be reduced to $O(r^2m)$. In this case, the *lcs* for $k = r$ needs to be recovered using ordinary *lcs* computations to connect the consecutive match points.

Remark: Space complexity can further be improved by applying a technique used in unconstrained *lcs* computation [3]. We can compute, instead of the entire *lcs* for each k , middle vertex $(n/2, j)$ (assume for simplicity that n is even) at which an *lcs* with respect to $P[1..k]$ passes. This can be done in $O(rm)$ space, and we can compute for all k the *lcs* length $L_{n/2,j,k}$ from vertex $(0, 0)$ to vertex $(n/2, j)$, and *lcs* length from $(n/2, j)$ to (n, m) . The latter is done in the reverse edit graph by calculating *lcs* from (n, m) to $(n/2, j)$, hence we denote it by $L_{n/2,j,l}^{\text{reverse}}$ for $0 \leq l \leq k$. Then for every k ,

$$\max_{j, 0 \leq l \leq k} L_{n/2,j,l} + L_{n/2,j,k-l}^{\text{reverse}}$$

is the *lcs* length for k , and it identifies a middle vertex. After the middle vertex $(n/2, j)$ on *lcs* for every k is found, the problem of finding the *lcs* from $(0, 0)$ to (n, m) can be solved in two parts: find the *lcs* from $(0, 0)$ to $(n/2, j)$, and find the *lcs* from $(n/2, j)$ to (n, m) for all k . These two subproblems can be solved recursively by finding the middle points. This way *lcs* can be obtained using $O(rm)$ space. The time complexity remains $O(rnm)$ because n is halved each time, and the area (in terms of number of vertices) covered in the edit graph is $O(nm)$, and at each vertex the total time spent is $O(r)$.

Next we propose a generalization of the constrained longest common subsequence problem. Given strings S_1, S_2 , and P , and a positive integer d the *edit distance*

constrained longest common subsequence problem for S_1 and S_2 with respect to string P , and distance d is to find a longest common subsequence lcs of S_1 and S_2 such that this lcs has a subsequence whose edit distance from P is smaller than d . Edit distance between two strings is the minimum number of edit operations required to transform one string to the other. The edit operations are insert, delete, and substitute.

Let $L_{i,j,k,t}$ be the length of an lcs for $S_1[1..i]$ and $S_2[1..j]$ such that the common subsequence contains a subsequence whose edit distance from $P[1..k]$ is exactly t .

Example: Suppose $S_1 = \text{bbaba}$, $S_2 = \text{abbaa}$ and $P = \text{ab}$. We have calculated before that the length of an lcs for S_1 and S_2 relative to P is 3. Thus $L_{5,5,2,0} = 3$. On the other hand the lcs bbaa of S_1 and S_2 contains the subsequence a , which is edit distance 1 away from P . Therefore $L_{5,5,2,1} = 4$.

We calculate all $L_{i,j,k,t}$ by a dynamic programming formulation. Optimal value of the edit distance constrained lcs problem is $\max_{0 \leq t < d} L_{n,m,r,t}$.

Theorem 2 For all i, j, k, t , $1 \leq i \leq n$, $1 \leq j \leq m$, $0 \leq k \leq r$, $0 \leq t < d$, $L_{i,j,k,t}$ satisfies

$$L_{i,j,k,t} = \max\{L'_{i,j,k,t}, L_{i,j-1,k,t}, L_{i-1,j,k,t}\} \quad (3)$$

where

$$L'_{i,j,k,t} = \max\{L''_{i,j,k,t}, L'''_{i,j,k,t}, L''''_{i,j,k,t}\} \quad (4)$$

where

$$L''_{i,j,k,t} = \begin{cases} 1 + L_{i-1,j-1,k-1,t} & \text{if } ((k = 1 \text{ and } t = 0) \text{ or} \\ & (k > 1 \text{ and } L_{i-1,j-1,k-1,t} > 0)) \\ & \text{and } S_1[i] = S_2[j] = P[k] \\ 0 & \text{otherwise} \end{cases}$$

$$L'''_{i,j,k,t} = \begin{cases} 1 + L_{i-1,j-1,0,0} & \text{if } (k = 0 \text{ and } t = 1) \text{ and } S_1[i] = S_2[j] \\ 1 + L_{i-1,j-1,k,t} & \text{else if } (k = 0 \text{ or } L_{i-1,j-1,k,t} > 0) \\ & \text{and } S_1[i] = S_2[j] \\ 0 & \text{otherwise} \end{cases}$$

$$L''''_{i,j,k,t} = \max\{D_{i,j,k,t}, X_{i,j,k,t}, I_{i,j,k,t}\} \quad (5)$$

where

$$D_{i,j,k,t} = \begin{cases} L_{i,j,k-1,t-1} & \text{if } t \geq 1 \\ 0 & \text{otherwise} \end{cases}$$

$$X_{i,j,k,t} = \begin{cases} L_{i,j,k-1,t-1} & \text{if } t \geq 1 \text{ and } S_1[i] = S_2[j] \\ 0 & \text{otherwise} \end{cases}$$

$$I_{i,j,k,t} = \begin{cases} L_{i,j,k,t-1} & \text{if } t \geq 1 \text{ and } S_1[i] = S_2[j] \\ 0 & \text{otherwise} \end{cases}$$

with boundary conditions $L_{i,0,k,0} = 0$, $L_{0,j,k,0} = 0$, for all i, j, k , $0 \leq i \leq n$, $0 \leq j \leq m$, $0 \leq k \leq r$.

Proof We claim that $L_{i,j,k,t}$ is the optimum length for any lcs for $S_1[1..i]$ and $S_2[1..j]$ such that the lcs contains a subsequence whose edit distance is t from $P[1..k]$. We prove the correctness of our formulation by induction on t for all i, j, k .

In the base case: when $t = 0$ the formulation becomes the same formulation as in Theorem 1, since now the lcs is required to contain P itself as a subsequence. Therefore, the correctness of this case follows from Theorem 1.

Assume that for $t - 1$ ($t \geq 1$), for all i, j, k , $L_{i,j,k,t-1}$ is the optimum length for any lcs for $S_1[1..i]$ and $S_2[1..j]$ such that the lcs contains a subsequence whose edit distance is t from $P[1..k]$. Consider the calculation of $L_{i,j,k,t}$ for all i, j, k when $t > 1$.

Our formulation uses the following observation: Let cs be a subsequence of an lcs of S_1 and S_2 . The minimum edit distance between cs and P can be calculated using insert, delete, and substitute operations in P , and using no operations in cs . To see this consider the edit operations between the symbols in cs , and in P . If an edit distance calculation deletes a symbol s in cs , we can instead insert the symbol s in P ; if a minimum edit distance calculation inserts a symbol s in cs , we can instead delete the symbol s in P ; and if a minimum edit distance calculation substitutes a symbol s' for s in cs , we can instead substitute a symbol s for s' in P to obtain the same edit distance.

We define an *edit path* at node (i, j) at distance t from $P[1..k]$ as a simple path from node $(0, 0)$ to node (i, j) , which includes a sequence of $l \geq 1$ distinct nodes $(a_1, b_1), (a_2, b_2), \dots, (a_l, b_l)$ such that the edit distance between the string $S_1[a_1]S_2[a_2] \dots S_1[a_l]$ ($= S_2[b_1] S_2[b_2] \dots S_2[b_l]$), and $P[1..k]$ is exactly t . We define $\#match$ on a given edit path to node (i, j) as the number of matching diagonal arcs on the path between the symbols in $S_1[1..i]$, and the symbols in $S_2[1..j]$, not necessarily involving matches in P . An optimal edit path at node (i, j) at distance t from $P[1..k]$ is an edit path at node (i, j) at distance t from $P[1..k]$ with maximum $\#match$. Thus $L_{i,j,k,t}$ is $\#match$ on an optimal edit path at node (i, j) at distance t from $P[1..k]$. In this case, $\#match = \#match(i, j, k, t)$ is a function of the indices i, j, k, t , but we omit these parameters when they are clear from the context.

We can extend any optimal edit path at node $(i, j - 1)$ at distance t from $P[1..k]$ with the horizontal arc $((i, j - 1), (i, j))$ to obtain an edit path at node (i, j) at distance t from $P[1..k]$. Such an extension does not change $\#match$ on the resulting edit path, and $L_{i,j,k,t} \geq L_{i,j-1,k,t}$.

Similarly we can extend any optimal edit path at node $(i - 1, j)$ at distance t from $P[1..k]$ with the vertical arc $((i - 1, j), (i, j))$ to obtain an edit path at node (i, j) at distance t from $P[1..k]$. This extension does not change $\#match$ on the resulting edit path, and $L_{i,j,k,t} \geq L_{i-1,j,k,t}$. Therefore, $L_{i,j,k,t} \geq \max\{L_{i,j-1,k,t}, L_{i-1,j,k,t}\}$.

By using a matching arc $((i - 1, j - 1), (i, j))$, we can obtain edit paths at node (i, j) at distance t from $P[1..k]$ by extending optimal edit paths at node $(i - 1, j - 1)$ at distance $t - 1$, or t from $P[1..k - 1]$, or $P[1..k]$.

First consider optimal edit paths at node $(i - 1, j - 1)$ at distance t from $P[1..k - 1]$. We will show that $L''_{i,j,k,t}$ stores the maximum $\#match$ obtained at node (i, j) by extending these edit paths.

If $S_1[i] = S_2[j] = P[k]$ then: We do not need to consider the case when $k = 1$ and $t = 0$ since $t = 0$ case is considered in the base case of the induction. If $k > 1$ and if there is an optimal edit path at node (i, j) at distance t from $P[1..k]$ (i.e. if

$L_{i-1,j-1,k-1,t} > 0$) then we can extend this edit path with a new match, and $\#match$ on the resulting edit path at node (i, j) at distance t from $P[1..k]$ becomes $L''_{i,j,k,t} = 1 + L_{i-1,j-1,k-1,t}$.

Next we consider optimal edit paths at node $(i-1, j-1)$ at distance t from $P[1..k]$. We will show that $L'''_{i,j,k,t}$ stores the maximum $\#match$ obtained at node (i, j) by extending these edit paths.

If $S_1[i] = S_2[j]$ then: If $k = 0$ and $t = 1$ then: We can extend an *lcs* path ending at node $(i-1, j-1)$ with respect to $P[1..k]$ with a match. In this case, $\#match$ in the resulting edit path is one more than $L_{i-1,j-1,0,0}$. Therefore, $L'''_{i,j,0,1} = 1 + L_{i-1,j-1,0,0}$. Otherwise if $k = 0$ then we can extend an optimal edit path at node $(i-1, j-1)$ at distance t from $P[1..k]$ with a match, and $\#match$ on the resulting edit path is $L'''_{i,j,k,t} = 1 + L_{i-1,j-1,k,t}$.

Any edit path at node (i, j) at distance $t-1$ from $P[1..k-1]$, or $P[1..k]$ can be modified by applying an edit operation in P . We can modify an edit path at node (i, j) at distance $t-1$ from $P[1..k-1]$ by deleting $P[k]$. Then on the resulting edit path $\#match$ remains the same, and the distance increases by 1. Therefore, we set $D_{i,j,k,t} = L_{i,j,k-1,t-1}$, and take it into account in $L''''_{i,j,k,t}$. We can modify an edit path at node (i, j) at distance $t-1$ from $P[1..k-1]$ by substituting $S_1[i] = S_2[j]$ for $P[k]$. Then on the resulting edit path $\#match$ remains the same, and the distance increases by 1. Therefore, we set $X_{i,j,k,t} = L_{i,j,k-1,t-1}$ if $S_1[i] = S_2[j]$, and take it into account in $L''''_{i,j,k,t}$. We can also modify an edit path at node (i, j) at distance $t-1$ from $P[1..k]$ by inserting $S_1[i] = S_2[j]$ in P after position k . Then on the resulting edit path $\#match$ remains the same, and the distance increases by 1. Therefore, we set $I_{i,j,k,t} = L_{i,j,k,t-1}$ if $S_1[i] = S_2[j]$, and take it into account in $L''''_{i,j,k,t}$. Combining all these $L''''_{i,j,k,t} = \max\{D_{i,j,k,t}, X_{i,j,k,t}, I_{i,j,k,t}\}$.

After setting $L'_{i,j,k,t} = \max\{L''_{i,j,k,t}, L'''_{i,j,k,t}, L''''_{i,j,k,t}\}$, $L'_{i,j,k,t}$ stores the maximum $\#match$ on edit paths at node (i, j) at distance t from $P[1..k]$ whose last arc is $((i-1, j-1), (i, j))$. If there is no such edit path then $L'_{i,j,k,t} = 0$.

From all possible optimal edit paths at neighboring nodes of (i, j) we can obtain their extensions ending at node (i, j) , and we can find an optimal edit path at node (i, j) at distance t from $P[1..k]$ for all k, t . We calculate, and store in $L_{i,j,k,t}$ maximum $\#match$ in such optimal edit paths. Considering the possible cases of the last arc on an optimal edit path at node (i, j) at distance t from $P[1..k]$ we also have $L_{i,j,k,t} \leq \max\{L'_{i,j,k,t}, L_{i,j-1,k,t}, L_{i-1,j,k,t}\}$. This concludes the proof of the theorem. •

All $L_{n,m,r,t}$ for $t = 0, 1, \dots, d-1$ can be computed in $O(dnrm)$ time, and using $O(drm)$ space using the formulation in Theorem 2 by noting that we only need rows $i-1$, and i during the calculations at row i . If an actual optimal edit path is desired then we can carry the edit path information for every k and t along with the calculations. This requires $O(dnrm)$ space since edit paths can be of length $O(n)$.

If we store match points (where the symbols of S_1 , S_2 , and P match) on these edit paths then we can reduce the required space to $O(dr^2m)$. In this case, the optimal edit path of the problem needs to be recovered using ordinary *lcs* computations to connect the consecutive match points.

Remark: Space complexity can further be improved by using the technique we used in our first algorithm. We can compute, instead of the entire edit path for each k , and t , a middle vertex $(n/2, j)$ (assume for simplicity that n is even) at which an edit path at distance t from $P[1..k]$ passes. This can be done in $O(drm)$ space, and we

can compute for all k , and t , $\#match$ $L_{n/2,j,l,u}$ on optimal edit path from vertex $(0, 0)$ to vertex $(n/2, j)$, and $\#match$ on optimal edit path from $(n/2, j)$ to (n, m) where $0 \leq \ell \leq k$, and $0 \leq u \leq t$. The latter, denoted by $L_{n/2,j,k-l,t-u}^{reverse}$, can be calculated in the reverse edit graph. Then for all k, t ,

$$\max_{j, 0 \leq \ell \leq k, 0 \leq u \leq t} L_{n/2,j,l,u} + L_{n/2,j,k-l,t-u}^{reverse}$$

is the optimum $\#match$ for k, t , and it identifies a middle vertex. After the middle vertex $(n/2, j)$ on optimal edit path for every k, t is found, the problem of finding an optimal edit path from $(0, 0)$ to (n, m) can be solved in two parts: find an optimal edit path from $(0, 0)$ to $(n/2, j)$, and find an optimal edit path from $(n/2, j)$ to (n, m) for all k, t . These two subproblems can be solved recursively. As a result an optimal edit path can be obtained using $O(drm)$ space. The time complexity remains $O(rnm)$ because n is halved each time, and the area (in terms of number of vertices) covered in the edit graph is $O(nm)$, and at each vertex the total time spent is $O(dr)$.

3 Conclusion

We have improved the time complexity of the constrained *lcs* problem from $O(rn^2m^2)$ to $O(rnm)$ where n , and m are the lengths of the given strings, and r is the pattern length. This improvement is achieved by a dynamic programming formulation which is different from what was proposed in [6]. In our formulation, the dimensions are ordered differently. We also extended the problem definition to use edit distances, and presented an $O(drn)$ time algorithm for the resulting edit distance constrained *lcs* problem.

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