



# Approximation hardness of deadline-TSP reoptimization<sup>☆</sup>

Hans-Joachim Böckenhauer<sup>a,\*</sup>, Joachim Kneis<sup>b,1</sup>, Joachim Kupke<sup>c,1</sup>

<sup>a</sup> Department of Computer Science, ETH Zurich, Switzerland

<sup>b</sup> Department of Computer Science, RWTH Aachen University, Germany

<sup>c</sup> Google Inc., Mountain View, USA

## ARTICLE INFO

### Article history:

Received 2 February 2007

Received in revised form 18 December 2008

Accepted 19 February 2009

Communicated by G. Ausiello

### Keywords:

Approximation algorithms

TSP with deadlines

Reoptimization

Inapproximability

## ABSTRACT

Given an instance of an optimization problem together with an optimal solution, we consider the scenario in which this instance is modified locally. In graph problems, e.g., a singular edge might be removed or added, or an edge weight might be varied, etc. For a problem  $U$  and such a local modification operation, let  $LM-U$  (local-modification- $U$ ) denote the resulting problem. The question is whether it is possible to exploit the additional knowledge of an optimal solution to the original instance or not, i.e., whether  $LM-U$  is computationally more tractable than  $U$ . While positive examples are known e.g. for metric TSP, we give some negative examples here: Metric TSP with deadlines (time windows), if a single deadline or the cost of a single edge is modified, exhibits the same lower bounds on the approximability in these local-modification versions as those currently known for the original problem.

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## 1. Introduction

Traditionally, optimization theory has been concerned with the task of finding good feasible solutions to (practically relevant) input instances, little or nothing about which is known in advance. Many applications, however, demand good, sometimes optimal, solutions to a limited set of input instances which reflect a supposedly-constant environment (imagine, e.g., an existing railway system or communications network). When this environment does change, maybe only slightly and maybe only locally, do we have no choice but to recompute some good feasible solution, effectively forgetting about the old one?

Here, we will analyze *local* modifications only. In a graph problem, for example, the cost of a single edge might change, an edge might be removed or added, or some other local parameter might be adjusted. Results related to this work pertain to the question by how much a given instance of an optimization problem may be varied if it is desired that optimal solutions to the original instance retain their optimality [8,9,11–13]. In contrast with this so-called “postoptimality analysis,” our approach here is to ask, if we cannot avoid losing the optimality of a given solution when an instance is varied arbitrarily, what can we do to *restore* the quality of a solution, maybe in an approximative sense?

Surely, for some problems, knowing an optimal solution to the original instance trivially makes their local-modification variants easy to approximate because the given optimal solution is itself a very good solution to the modified instance. For

<sup>☆</sup> This work was partially supported by SNF grant 200021-109252/1. An extended abstract of this work has been presented at IFIP TCS 2006 [H.-J. Böckenhauer, L. Forlizzi, J. Hromkovič, J. Kneis, J. Kupke, G. Proietti, P. Widmayer, Reusing optimal TSP solutions for locally modified input instances (extended abstract), in: Proc. of the 4th IFIP International Conference on Theoretical Computer Science, Springer-Verlag, 2006, pp. 251–270].

\* Corresponding address: Department of Computer Science, ETH Zurich, Universitätsstr. 6, CH-8092 Zurich, Switzerland. Tel.: +41 44 6328183; fax: +41 44 6321390.

E-mail addresses: [hjb@inf.ethz.ch](mailto:hjb@inf.ethz.ch) (H.-J. Böckenhauer), [kneis@cs.rwth-aachen.de](mailto:kneis@cs.rwth-aachen.de) (J. Kneis), [joachimk@google.com](mailto:joachimk@google.com) (J. Kupke).

<sup>1</sup> This author was staying at ETH Zurich when this work was done.

example, adding an edge in the instance of a coloring problem will increase the cost of an optimal solution by at most the amount of one – an excellent approximation, but certainly not the object of our interest.

This concept of local modification was investigated for metric TSP in [3], where modifications of edge costs were considered, and in [1,2], where the local modification consisted of vertex deletions and vertex insertions.

TSP with time windows is one of the fundamental problems in operations research [7]. In this generalization of TSP, certain vertices must be reached within a certain time window. Usually, only heuristic algorithms are used to attack it although the question how hard it is w. r. t. approximability has only recently been resolved in [5,4], where even an  $\Omega(n)$  lower bound on the polynomial-time approximability of metric TSP ( $\Delta$ -TSP for short) with time windows was shown, in contrast to the constant approximability of  $\Delta$ -TSP. This lower bound already holds for the special case of this problem where all time windows are immediately open. This special case we will call metric TSP with deadlines, or  $\Delta$ -DLTSP for short. Here, we consider local-modification versions of  $\Delta$ -DLTSP. We show that, already if we only allow a single deadline to be changed, and only by an amount of one time unit, the resulting problem, LM(D)- $\Delta$ -DLTSP, has the same lower bound on the approximation ratio as  $\Delta$ -DLTSP, namely that of asymptotically  $n/2$ . Let us underscore the importance of this negative result: Not only does TSP with deadlines remain an intractable problem in its local-modification version, but the extra knowledge of an optimal solution to a related instance does not even help a single bit. Likewise, we will establish a lower bound of  $(2 - \varepsilon)$ , for any  $\varepsilon > 0$ , for LM(D)- $\Delta$ -DLTSP with a constant number of deadlines, the same bound as is known for  $\Delta$ -DLTSP with a constant number of deadlines [5,4]. These results can also be obtained if we modify the cost of an edge rather than a deadline. Note that the original problem without reoptimization was proven to be approximable within a factor of 2.5 in [5,4].

So, while additional information about an optimal solution to a related input instance may be useful to some extent as shown e. g. in [3], the local-modification problem variant may remain exactly as hard as the original problem. Very recently, in [6] the reoptimization of  $\Delta$ -DLTSP according to other local modifications such as addition and deletion of vertices was investigated.

The paper is subdivided into four main sections. In Section 2, we will formally define TSP with deadlines as well as its local-modification version. Section 3 is devoted to inapproximability results for the local-modification version of TSP with deadlines where there is a bounded number of deadline vertices, and Section 4 will deal with the general problem, where there may be an unbounded number of deadline vertices. In these two sections, we deal with the case where the local modification consists of changing a deadline. Finally, in Section 5, we analyze the case where the local modification consists of changing the costs of an edge.

## 2. Preliminaries

To begin with, let us define TSP with deadlines formally.

**Definition 2.1.** Let  $G = (V, E)$  be a complete graph weighted by  $c: E \rightarrow \mathbb{N}^+$ . We call  $(s, D, d)$  a *deadline set* for  $G$  if  $s \in V$ ,  $D \subseteq V \setminus \{s\}$ , and  $d: D \rightarrow \mathbb{N}^+$ . A vertex  $v \in D$  is called *deadline vertex*. A path  $(v_0, v_1, \dots, v_n)$  *satisfies the deadlines* iff  $s = v_0$  and, for all  $v_i \in D$ , we have  $\sum_{j=1}^i c(\{v_{j-1}, v_j\}) \leq d(v_i)$ .

A cycle  $(v_0, v_1, \dots, v_n, v_0)$  *satisfies the deadlines* iff it contains a path  $(v_0, v_1, \dots, v_n)$  satisfying the deadlines.

**Definition 2.2.** The problem  $\Delta$ -DLTSP is defined as follows:

*Input:* A complete weighted graph  $G = (V, E, c)$  with edge costs  $c: E \rightarrow \mathbb{N}^+$  satisfying the  $\Delta$ -inequality, a deadline set  $(s, D, d)$  for  $G$ , and a Hamiltonian cycle (of arbitrary cost) satisfying the deadlines.<sup>1</sup>

*Problem:* Find a minimum-cost Hamiltonian cycle satisfying all deadlines.

If  $|D|$  is a constant  $k$ , the resulting subproblem is  $k$ - $\Delta$ -DLTSP.

For TSP with deadlines, at least two possible local modifications arise. Firstly, we can change the deadline of a deadline vertex, which leads to the following problem:

**Definition 2.3.** The optimization problem LM(D)-DLTSP is defined as follows:

*Input:* A complete weighted graph  $G = (V, E, c)$ , a deadline set  $O = (s, D, d_O)$  for  $G$  with a minimal Hamiltonian cycle satisfying the deadlines  $O$ , a new deadline set  $N = (s, D, d_N)$  such that  $d_O$  and  $d_N$  differ in exactly one vertex, and a Hamiltonian cycle (of arbitrary cost) satisfying  $N$ .

*Problem:* Find a minimum-cost Hamiltonian cycle satisfying  $N$ .

By LM(D)- $k$ -DLTSP, LM(D)- $\Delta$ -DLTSP and LM(D)- $k$ - $\Delta$ -DLTSP, we denote the canonical special cases of LM(D)-DLTSP.

The second local modification we consider here, is a change in the cost function for one edge.

<sup>1</sup> Requiring a feasible Hamiltonian cycle as part of the input ensures that the problem is in *NPO*. Otherwise, it would even be a hard problem to find a feasible solution. For details, see [5,4].

**Definition 2.4.** The optimization problem  $\text{LM}(E)\text{-DLTSP}$  is defined as follows:

*Input:* A complete weighted graph  $G_O = (V, E, c_O)$ , a deadline set  $(s, D, d)$  for  $G$ , with a minimum-cost Hamiltonian cycle satisfying the deadlines, and a complete weighted graph  $G_N = (V, E, c_N)$  such that  $c_O$  and  $c_N$  differ in exactly one edge, and a Hamiltonian cycle (of arbitrary cost) in  $G_N$  satisfying the deadlines.

*Problem:* Find a minimum-cost Hamiltonian cycle in  $G_N$  satisfying the deadlines.

By  $\text{LM}(E)\text{-}k\text{-DLTSP}$ ,  $\text{LM}(E)\text{-}\Delta\text{-DLTSP}$  and  $\text{LM}(E)\text{-}k\text{-}\Delta\text{-DLTSP}$ , we denote the canonical special cases of  $\text{LM}(E)\text{-DLTSP}$ .

For our proofs, we will need some reductions from the following two problems, which can easily be shown to be *NP*-hard analogously to the proof of the *NP*-hardness of the *restricted Hamiltonian cycle problem*, as presented, e.g., in [10].

**Definition 2.5.** The *restricted Hamiltonian path problem*, RHP for short, is the following decision problem:

*Input:* A graph  $G = (V, E)$ , two different vertices  $s, t \in V$ , and a given Hamiltonian path  $P$  from  $s$  to  $t$ .

*Output:* YES if  $G$  contains a Hamiltonian path starting in  $s$ , but ending in some vertex  $v \neq t$ , No otherwise.

**Definition 2.6.** The *restricted Hamiltonian path problem with fixed endpoint*, FRHP for short, is the following decision problem:

*Input:* A graph  $G = (V, E)$ , three pairwise different vertices  $s, t, x \in V$ , and a given Hamiltonian path  $P$  from  $s$  to  $t$ .

*Output:* YES if  $G$  contains a Hamiltonian path starting in  $s$  and ending in  $x$ , No otherwise.

### 3. Bounded number of deadline vertices

We start with the case where only few deadline vertices occur. Note that  $k\text{-}\Delta\text{-DLTSP}$  can be approximated within a ratio of 2.5 [5,4]. Furthermore, a lower bound of  $2 - \varepsilon$  on the approximability, for every small  $\varepsilon > 0$ , can be proved [5,4]. We will show that this lower bound also holds for  $\text{LM}(D)\text{-}k\text{-}\Delta\text{-DLTSP}$ , even for  $k = 4$ .

For the proof we will distinguish two subproblems, namely increasing and decreasing the deadline of one deadline vertex. We will start with the increasing case.

**Lemma 3.1.** *Let  $\varepsilon > 0$ . There is no polynomial-time  $(2 - \varepsilon)$ -approximation algorithm for the subproblem of  $\text{LM}(D)\text{-}2\text{-}\Delta\text{-DLTSP}$  where one deadline is increased by  $\xi$  time units,  $1 \leq \xi < n$ , unless  $P = \text{NP}$ .*

**Proof.** By means of a reduction, we will show that such an approximation algorithm could be used to solve FRHP. Let  $\varepsilon > 0$ .

Let  $(G', P)$  be an input instance for FRHP where  $G' = (V', E')$ ,  $|V'| = n + 1$ ,  $s', t', x \in V'$ , and  $P$  is a Hamiltonian path from  $s'$  to  $t'$ . Pick a  $\gamma > \frac{5n+3}{2\varepsilon}$  (which implies  $\frac{4\gamma+n+1}{2\gamma+3n+1} > 2 - \varepsilon$ ).

We construct a complete weighted graph  $G = (V, E, c)$  as part of an input for  $\text{LM}(D)\text{-}2\text{-}\Delta\text{-DLTSP}$  as shown in Fig. 1: We set  $V := V' \dot{\cup} \{s, D_1, D_2\}$ , and, for any edge  $e$  between two vertices  $v_1, v_2 \in V'$ , let  $c(e) = 1$  if  $e \in E'$  and  $c(e) = 2$  otherwise. All edges depicted in Fig. 1 have the indicated costs while non-depicted edges obtain maximal possible costs that do not violate the triangle inequality. In particular, all vertices  $v' \in G' \setminus \{s', t', x\}$  are connected to  $D_1, D_2$ , and  $s$  in the same way as the depicted vertex  $v$ . We set the deadlines  $d(D_1) = \gamma + n$  and  $d(D_2) = 2\gamma + n + 1$ .

For these deadlines, one optimal solution  $\bar{C}$  is the cycle  $s, D_1, D_2, t', \dots, s'$ ,  $s$ , which uses the Hamiltonian path  $P$  backwards from  $t'$  to  $s'$  in  $G'$ . It costs exactly  $\gamma + 1 + \gamma + \gamma + n + \gamma = 4\gamma + n + 1$ .

Visiting  $D_1$  and all vertices in  $G'$  before  $D_2$  is impossible since it would violate the deadline at  $D_2$  or  $D_1$  in any case: On the one hand, any solution starting with  $s, D_1, v'$ , for some  $v' \in G'$ , using some path in  $G'$  from  $v'$  to  $t'$  and then proceeding to  $D_2$  can only satisfy the deadline at  $D_2$  if it leaves out some vertex in  $G'$ . On the other hand, any solution starting with  $s, s'$  and some partial tour to  $x$  and then visiting  $D_1$  before returning to  $G'$ , also has to leave out some vertex from  $G'$  in order to obey the deadline at  $D_2$ . This is due to the fact that including  $D_1$  in some tour through  $G'$  incurs an extra cost of at least 2. Visiting  $l$  still unvisited vertices in  $G'$  after  $D_2$  incurs an additional cost of at least  $2\gamma + l$ ; thus, a solution of this type cannot improve over  $\bar{C}$ .

Now, we increase  $d(D_1)$  by  $\xi$ . If  $G'$  contains a Hamiltonian path  $P'$  from  $s'$  to  $x$ , a new optimal solution is  $s, P', D_1, D_2, s$ , and it costs  $\gamma + n + 1 + \gamma + 2n = 2\gamma + 3n + 1$ . If  $G'$  does not contain such a path, it is not possible to visit all vertices in  $G'$  before reaching  $D_1$  and  $D_2$ . As  $c(\{t', D_1\}) \geq 2$ , we cannot follow the given Hamiltonian path  $P$  because this would violate the deadline  $d(D_2)$ . As in the old instance, including  $D_1$  in some Hamiltonian path in  $G'$  would violate the deadline at  $D_2$ . Similar arguments hold for every other possibility. Hence,  $\bar{C}$  remains an optimal solution in this case. Thus, we could use any approximation algorithm with an approximation guarantee better than

$$\frac{4\gamma + n + 1}{2\gamma + 3n + 1} > 2 - \varepsilon$$

to solve FRHP. This is why approximating this subproblem of  $\text{LM}(D)\text{-}2\text{-}\Delta\text{-DLTSP}$  within  $2 - \varepsilon$  is *NP*-hard.  $\square$

**Lemma 3.2.** *Let  $\varepsilon > 0$ . There is no polynomial-time  $(2 - \varepsilon)$ -approximation algorithm for the subproblem of  $\text{LM}(D)\text{-}4\text{-}\Delta\text{-DLTSP}$  where one deadline is decreased by  $\xi$  time units,  $1 \leq \xi < n$ , unless  $P = \text{NP}$ .*

**Proof.** Let  $\varepsilon > 0$ . For the proof, we will use a reduction from RHP.

Let  $(G', P)$  be an input instance for RHP where  $G' = (V', E')$ ,  $|V'| = n + 1$ ,  $s', t' \in V'$ , and  $P$  is a Hamiltonian path from  $s'$  to  $t'$ . Pick some  $\gamma$  such that  $\frac{4\gamma}{2\gamma+8n} > 2 - \varepsilon$ .

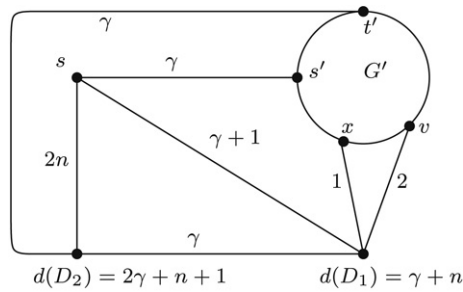


Fig. 1. Increasing a deadline. All vertices  $v' \in V' \setminus \{s', t', x\}$  are connected like  $v$ .

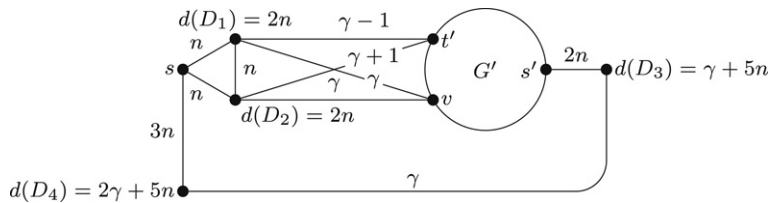


Fig. 2. Decreasing a deadline. All vertices  $v' \in V' \setminus \{s', t'\}$  are connected like  $v$ .

We construct a complete weighted graph  $G = (V, E, c)$  as part of an input for LM(D)-4- $\Delta$ -DLTSP as shown in Fig. 2: We set  $V := V' \cup \{s, D_1, D_2, D_3, D_4\}$ , and, for any edge  $e$  between two vertices  $v_1, v_2 \in V'$ , let  $c(e) = 1$  if  $e \in E'$  and  $c(e) = 2$  otherwise. All edges depicted in Fig. 2 have the indicated costs while non-depicted edges obtain maximal possible costs not violating the triangle inequality.

The initial deadlines are depicted in Fig. 2. In this setting, an optimal solution is the cycle  $s, D_2, D_1, t', \dots, s', D_3, D_4, s$ , which contains the Hamiltonian path  $P$  from  $s'$  to  $t'$ . This path costs  $2n + \gamma - 1$  on its way to  $G'$ , spends  $n$  on the path from  $t'$  to  $s'$ , and reaches  $s$  at time  $2\gamma + 8n - 1$ . Note that visiting  $D_1$  before  $D_2$  leads to an additional cost of at least 1 and is thus not optimal.

Now, we decrease the deadline  $d(D_1)$  by  $\xi$ , whereby the old optimal solution becomes infeasible. Any new solution must visit  $D_1$  before  $D_2$ . If we try to reuse the Hamiltonian path from  $t'$  to  $s'$ , we have to spend  $2n + \gamma + 1$  on the way to  $t'$ . Therefore, we cannot reach  $D_3$  if we follow the complete Hamiltonian path. Furthermore, we cannot visit any vertex  $v \in V'$  between visiting  $D_3$  and  $D_4$  because  $D_3$  is not reached before  $4n + \gamma$ , going back to  $V'$  would cost another  $2n$ , and the cheapest path from  $V'$  to  $D_4$  costs more than  $\gamma$ . This is why any solution using a Hamiltonian path between  $s'$  and  $t'$  violates one of the deadlines  $d(D_3), d(D_4)$ .

If  $G'$  contains a Hamiltonian path  $P$  from  $s'$  to some  $v \neq t'$ , the new optimal solution contains this path in reverse on its way to  $D_3$ . The path  $s, D_1, D_2, P, D_3, D_4$  visits all vertices in  $V'$  between  $v$  and  $s'$  and reaches  $D_3$  at time  $\gamma + 5n$ . Therefore, this new optimal solution costs  $2\gamma + 8n$ .

If  $G'$  does not contain such a Hamiltonian path, the optimal solution cannot visit all vertices in  $V'$  before reaching  $D_3$  or even  $D_4$ , and consequently, it is more expensive than  $4\gamma$ . Thus, we could use an approximation algorithm with an approximation guarantee better than

$$\frac{4\gamma}{2\gamma + 8n} > 2 - \varepsilon$$

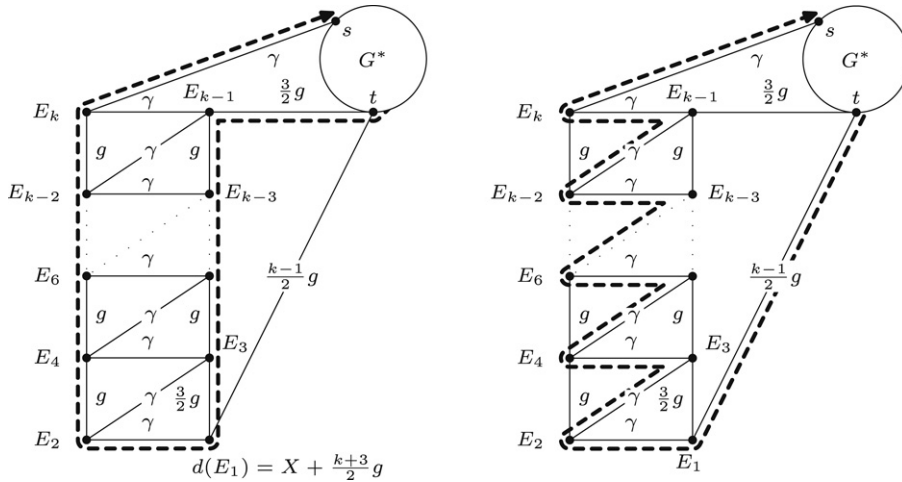
to solve RHP. Hence, approximating this subproblem of LM(D)-4- $\Delta$ -DLTSP within  $2 - \varepsilon$  is NP-hard.  $\square$

Lemmas 3.1 and 3.2 directly imply our main result of this section.

**Theorem 3.3.** *There is no polynomial-time  $(2 - \varepsilon)$ -approximation algorithm for LM(D)-4- $\Delta$ -DLTSP, unless  $P = NP$ .  $\square$*

#### 4. Unbounded number of deadline vertices

When the number of deadline vertices is unbounded, we are able to show a linear lower bound on the approximability of LM(D)- $\Delta$ -DLTSP. Our reduction from RHP involves two steps. A first construction will guarantee that an optimal path becomes shorter by a constant factor if a Hamiltonian path exists in the RHP instance. A second construction inflates this advantage. Tours which start at time  $X$ , in contrast to those that start between times  $X + g$  and  $X + 2g$ , may spend some extra time to visit a group of vertices which, unless visited early, will cause belated tours to run  $k$  times zigzag across a huge distance  $\gamma$ . Note that the construction will be such that tours starting after  $X + 2g$  will violate some deadline in any case.



**Fig. 3.** The zigzag construction for the proof of Lemma 4.1. The left-hand side shows the optimal path if  $t$  is reached at time  $X$ . The right-hand side shows the optimal solution if  $t$  is reached after  $X + g$ . We set  $d(E_{i+1}) := d(E_i) + \gamma$ .

The following lemma describes the construction in detail. See Fig. 3 for an overview.

**Lemma 4.1 (Zigzag Lemma).** *Let  $X, g, k, \gamma \in \mathbb{N}$  such that  $k$  is even and  $\gamma \geq g$ . Let  $G^* = (V^*, E^*)$  be a complete graph with deadline set  $(s, D^*, d^*)$  such that any Hamiltonian path in  $G^*$  respecting the deadlines ends in the same vertex  $t$ . Then, we can construct a complete graph  $G \supset G^*$  and deadlines  $(s, D, d)$  such that  $D \supset D^*$ ,  $d_{D^*} = d^*$  and any path  $P$  in  $G^*$  that reaches  $t$  in time  $X$  can be extended to a Hamiltonian cycle  $C = PP's$ , where  $P'$  is a path visiting all vertices outside  $G^*$ , which costs at most*

$$X + kg + 2\gamma,$$

while any path that reaches  $t$  after  $X + g$ , but before  $X + 2g$  can only be extended to a Hamiltonian cycle which costs at least

$$X + \frac{k+1}{2}g + k\gamma.$$

**Proof.** We construct  $G = (V, E)$  with  $V = V^* \cup \{E_1, \dots, E_k\}$  and edge costs as depicted in Fig. 3. To all other edges, we assign maximal possible costs that do not violate the triangle inequality. Note that the edge  $\{t, E_1\}$  costs exactly the same as the path  $E_{k-1}, E_{k-3}, \dots, E_1$ .

We set the deadlines

$$d(E_1) := X + \frac{k+3}{2}g \quad \text{and}$$

$$d(E_{i+1}) := d(E_i) + \gamma \quad \text{for all } i \in \{1, \dots, k-1\}.$$

If a path reaches  $t$  strictly after  $X + g$ , it must proceed immediately to  $E_1$ . Note that it cannot use any other edge since it would have to use an edge of an additional cost of at least  $\frac{3}{2}g$ , then. Together with even the shortest path to  $E_1$ , this would violate this deadline. But then, it is forced to follow the sequence  $E_2, E_3, \dots, E_k$  to reach every deadline since even if it visited  $E_3$  before  $E_2$ , it would incur an extra cost of  $\frac{3}{2}g$ , and this would violate the deadline of  $E_2$ . Analogously, proceeding from  $E_1, E_2$  directly to  $E_4$  would necessarily violate the deadline of  $E_3$ . Hence, the Hamiltonian cycle costs strictly more than  $X + g + \frac{k-1}{2}g + k\gamma$ .

A path that visits  $t$  before time  $X$  can visit  $E_{k-1}, E_{k-3}, \dots, E_3$  before  $E_1$  because this path to  $E_1$  costs at most

$$X + \frac{3}{2}g + \left(\frac{k}{2} - 2\right)g + \frac{3}{2}g = X + \frac{k+2}{2}g \leq d(E_1).$$

Closing the cycle to  $s$ , we obtain a cost of at most

$$X + \frac{k+2}{2}g + \left(\frac{k}{2} - 1\right)g + 2\gamma = X + kg + 2\gamma. \quad \square$$

We will now employ Lemma 4.1 to prove the desired lower bound. Again, we will consider the cases of increasing and decreasing a deadline separately.

**Lemma 4.2.** *Let  $\varepsilon > 0$ . There is no polynomial-time  $\left(\left(\frac{1}{2} - \varepsilon\right) \cdot |V|\right)$ -approximation algorithm for the subproblem of LM(D)- $\Delta$ -DLTSP where one deadline is increased by  $\xi$  for some  $1 \leq \xi < n$ , unless  $P = NP$ .*

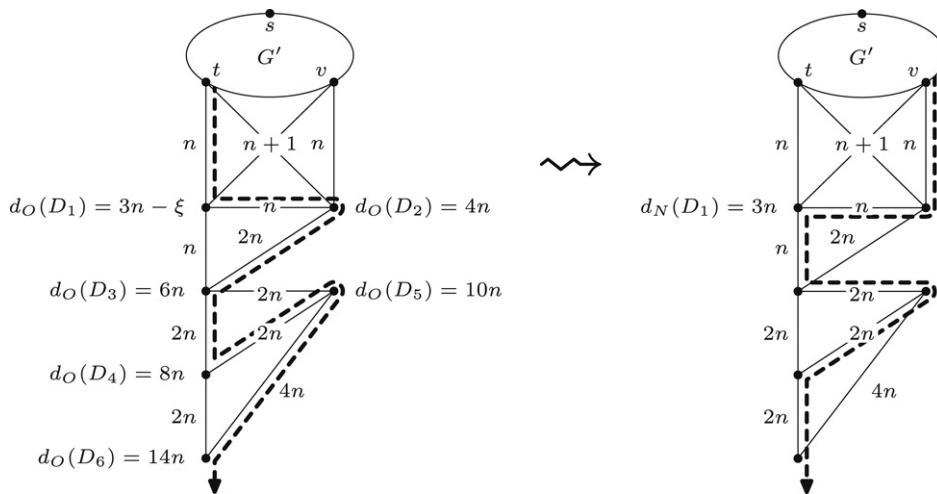


Fig. 4. Increasing a deadline: If the deadline for the vertex  $D_1$  is increased, using a Hamiltonian path from  $s$  to  $v$  leads to a new optimal solution.

**Proof.** By means of a reduction, we will show that such an approximation algorithm could be used to solve RHP.

Let  $(G', P)$  be an input instance for RHP, where  $G' = (V', E')$ ,  $|V'| = n + 1$ ,  $s, t \in V'$ , and  $P$  is a Hamiltonian path from  $s$  to  $t$ .

We construct a complete weighted graph  $G = (V, E, c)$  from  $G'$  as part of the input for the  $\text{LM}(\text{D})$ - $\Delta$ -DLTSP instance as follows. The graph  $G$  contains a subgraph  $G^*$  as shown in Fig. 4 which is then complemented by the zigzag construction from Lemma 4.1.

We set  $V^* = V' \cup \{D_1, \dots, D_6\}$  and, for any edge  $e$  between two vertices  $v_1, v_2 \in V'$ ,  $c(e) = 1$ , if  $e \in E'$ , and  $c(e) = 2$  otherwise. To the other edges, we assign costs as depicted in Fig. 4, and maximal possible costs not violating the triangle inequality to the non-depicted edges; all edges from non-depicted vertices from  $G'$  to  $D_1$  and  $D_2$  get the same costs as the edges from  $v$ . We set the deadlines  $d_O(D_i)$  according to Fig. 4. In particular, the vertex  $D_1$  assumes a deadline of  $3n - \xi$  for some  $1 \leq \xi < n$ . Additionally, all vertices in  $G'$  assume a deadline of  $2n$ .

To construct the graph  $G_O$  of our desired  $\text{LM}(\text{D})$ - $\Delta$ -DLTSP instance from  $G^*$ , we use the zigzag construction defined in Lemma 4.1 with parameters  $X = 10n$ ,  $g = 2n$ ,  $k \geq (n + 7) \frac{1-\varepsilon}{\varepsilon}$ , and  $\gamma \geq \frac{2kn+10n}{\varepsilon}$ . This guarantees  $2kn + 10n \leq \varepsilon\gamma$  and  $k \geq (1 - \varepsilon)(k + n + 7) = (1 - \varepsilon) \cdot |V|$ .

The given optimal Hamiltonian tour  $\bar{C}$  in  $G$  starts in  $s$ , uses the given Hamiltonian path  $P$  in  $G'$  to  $t$ , and afterwards follows the sequence  $D_1, D_2, D_3, D_4, D_5, D_6$  before entering the zigzag part of  $G$ . Hence, it reaches  $D_6$  in time  $13n$ . It is easy to see that  $\bar{C}$  is indeed optimal for the old deadline set: Any tour visiting  $D_2$  before  $D_1$  will violate the deadline at  $D_1$ . Following the zigzag construction, this leads to an overall cost of at least  $10n + \frac{k+1}{2}g + k\gamma$ .

Note that any path alternating between the zigzag part of  $G$  and  $G^*$  will necessarily violate the deadline of some yet-to-be-visited vertex  $E_i$  for some  $i \in \{2, \dots, k\}$ .

In  $G_N$ , we change the deadline for  $D_1$  from  $3n - \xi$  to  $d_N(D_1) = 3n$ . The tour  $\bar{C}$  remains a feasible solution. If  $G'$  contains a Hamiltonian path from  $s$  to some vertex  $v \neq t$ , an optimal solution uses this path and follows the sequence  $D_2, D_1, D_3, D_5, D_4, D_6$ . This solution reaches  $D_6$  in time  $10n$ . By Lemma 4.1, this cycle has overall costs  $10n + kg + 2\gamma$ .

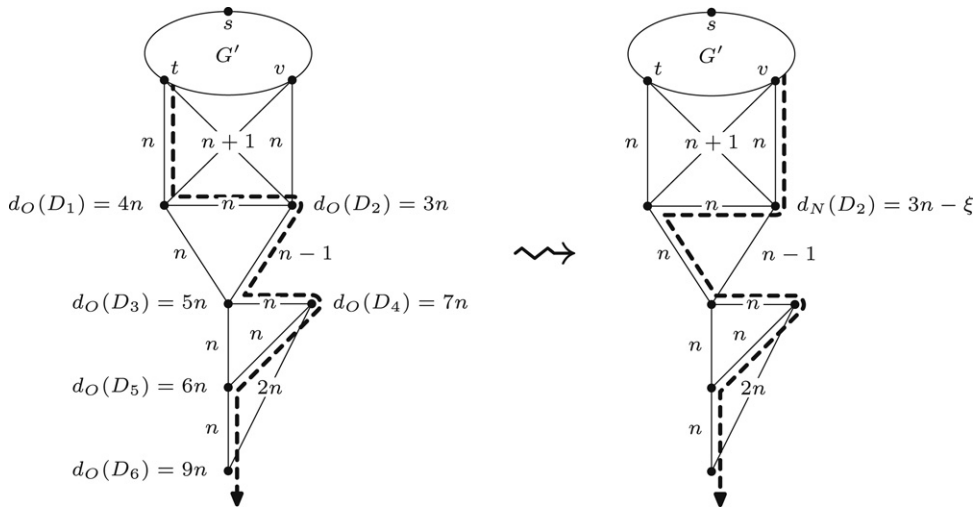
If  $G'$  does not contain any Hamiltonian path to such a vertex  $v$ , then  $\bar{C}$  remains the optimal solution. By Lemma 4.1, we again obtain a cost of  $10n + \frac{k+1}{2}g + k\gamma$ . Together with our choices of the parameters, this leads to a ratio of at least

$$\begin{aligned} \frac{10n + \frac{k+1}{2} \cdot 2n + k\gamma}{10n + k \cdot 2n + 2\gamma} &> \frac{k\gamma}{2kn + 10n + 2\gamma} \\ &> \frac{k\gamma}{(2 + \varepsilon)\gamma} = \frac{k}{2 + \varepsilon} \geq \frac{1 - \varepsilon}{2 + \varepsilon} (k + n + 7) \geq \left(\frac{1}{2} - \varepsilon\right) |V|. \end{aligned}$$

Hence, a polynomial-time  $(\frac{1}{2} - \varepsilon)|V|$ -approximation algorithm could be used to solve RHP.  $\square$

**Lemma 4.3.** Let  $\varepsilon > 0$ . There is no polynomial-time  $((\frac{1}{2} - \varepsilon)|V|)$ -approximation algorithm for the subproblem of  $\text{LM}(\text{D})$ - $\Delta$ -DLTSP where one deadline is decreased by  $\xi \geq 1$  unless  $P = NP$ .

**Proof.** By means of a reduction, we will show that such an approximation algorithm could be used to solve the RHP. Let  $(G', P)$  be an input instance for RHP, where  $G' = (V', E')$ ,  $|V'| = n + 1$ ,  $s, t \in V'$ , and  $P$  is a Hamiltonian path from  $s$  to  $t$ . We construct a complete weighted graph  $G = (V, E, c)$  as part of an input for the  $\text{LM}(\text{D})$ - $\Delta$ -DLTSP as follows. The graph  $G$  contains a subgraph  $G^*$  as shown in Fig. 5 which is then complemented by the zigzag construction from Lemma 4.1.



**Fig. 5.** Decreasing a deadline: If the deadline for the vertex  $D_2$  is decreased, the old optimal solution (depicted on the left side) is not feasible anymore. If  $G'$  contains a Hamiltonian path from  $s$  to  $v$ , we obtain the depicted new optimal solution. If no such Hamiltonian path exists, the new optimal solution must follow  $D_2, D_1, D_3, D_5, D_4, D_6$ .

We set  $V^* = V' \cup \{D_1, \dots, D_6\}$  and, for any edge  $e$  between two vertices  $v_1, v_2 \in V'$ ,  $c(e) = 1$ , if  $e \in E'$ , and  $c(e) = 2$  otherwise. For the other edges, we assign the costs as depicted in Fig. 5, and maximal possible costs not violating the triangle inequality to the non-depicted edges, and we set the deadlines  $d_O(D_i)$  according to Fig. 5.

Again, we use Lemma 4.1 to extend the graph  $G^*$ , using the parameters  $X = 7n - 1$ ,  $g = n$ ,  $k \geq (n + 7) \frac{1-\varepsilon}{\varepsilon}$ , and  $\gamma \geq \frac{(k+7)n}{\varepsilon}$ . The given optimal solution  $\bar{C}$  uses the Hamiltonian path from  $s$  to  $t$  and visits  $D_1, D_2, D_3, D_4, D_5, D_6$  afterwards.  $D_6$  is reached at time  $7n - 1$ , therefore the optimal Hamiltonian cycle  $\bar{C}$  costs at most  $7n - 1 + kn + 2\gamma$ .

Note that any path alternating between the zigzag part of  $G$  and  $G^*$  will necessarily violate the deadline of some yet-to-be-visited vertex  $E_i$  for some  $i \in \{2, \dots, k\}$ .

In  $G_N$ , we change the deadline for  $D_2$  as follows:  $d_N(D_2) = 3n - \xi$  for some  $1 \leq \xi < n$ . The cycle  $\bar{C}$  is no feasible solution anymore since it visits  $D_2$  at time  $3n$ . If  $G'$  contains a Hamiltonian path from  $s$  to some vertex  $v \neq t$ , an optimal solution for  $G_N$  uses this path and follows the sequence  $D_2, D_1, D_3, D_4, D_5, D_6$ . It reaches  $D_7$  in time  $7n$ . Hence, by Lemma 4.1, extending it to a Hamiltonian cycle costs at most

$$7n + kn + 2\gamma.$$

If  $G'$  does not contain a Hamiltonian path to any vertex  $v \neq t$ , an optimal solution will visit all vertices in  $V'$  and then  $D_2$ . This vertex is not reached before  $2n + 1$ . Thus, following the cheap path  $D_2, D_1, D_3, D_4, D_5$ , the deadline of vertex  $D_5$  is violated. Hence the solution must visit the remaining vertices in the order given by  $D_1, D_3, D_5, D_4, D_6$ . It does not reach  $D_6$  before  $8n + 1$ , hence by Lemma 4.1 the corresponding Hamiltonian cycle costs at least

$$7n + \frac{k+1}{2}n + k\gamma.$$

We obtain the ratio

$$\begin{aligned} \frac{7n + \frac{k+1}{2}n + k\gamma}{7n + kn + 2\gamma} &> \frac{k\gamma}{(7+k)n + 2\gamma} \geq \frac{k\gamma}{(2+\varepsilon)\gamma} = \frac{k}{2+\varepsilon} \\ &\geq \frac{1-\varepsilon}{2+\varepsilon}(k+n+7) = \frac{1-\varepsilon}{2+\varepsilon}|V| \geq \left(\frac{1}{2} - \varepsilon\right)|V|. \end{aligned}$$

Hence, a polynomial-time  $(\frac{1}{2} - \varepsilon)|V|$ -approximation algorithm could be used to solve RHP.  $\square$

From Lemmas 4.2 and 4.3 we can directly conclude the following:

**Theorem 4.4.** Let  $\varepsilon > 0$ . There is no polynomial-time  $((\frac{1}{2} - \varepsilon)|V|)$ -approximation algorithm for LM(D)- $\Delta$ -DLTSP, unless  $P = NP$ .  $\square$

### 5. Modifying edge costs

In the previous sections, we have analyzed deadline modifications. In what follows, we will show similar results for local modifications of edge costs. We start with the case of TSP with only a bounded number of deadlines.

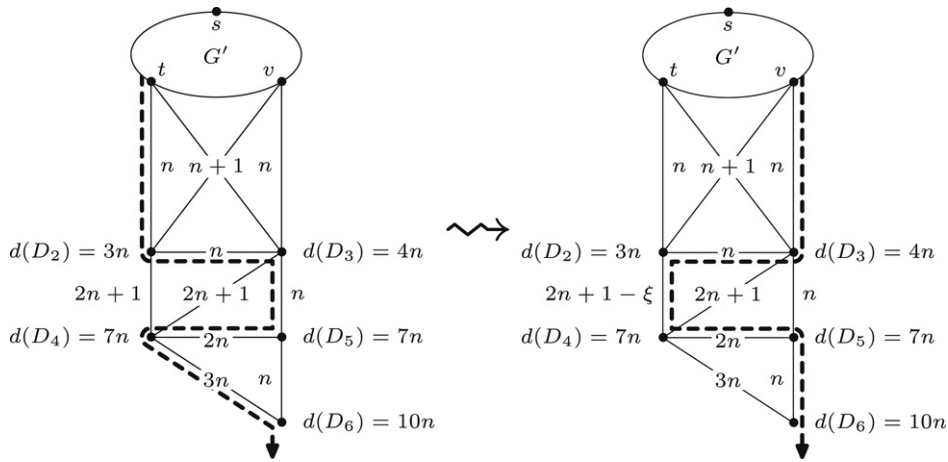


Fig. 6. Decreasing edge costs: If the cost of the edge  $\{D_2, D_4\}$  is decreased, using a Hamiltonian path from  $s$  to  $v$  leads to a new optimal solution.

**Theorem 5.1.** *Let  $\varepsilon > 0$ . There is no polynomial-time  $(2 - \varepsilon)$ -approximation algorithm for LM(E)-2- $\Delta$ -DLTSP even if only the cost of one edge is changed, unless  $P = NP$ .*

**Proof.** Let  $1 \leq \alpha < n$ . In the case of increasing edge costs, we use the same construction as in Fig. 2. Decreasing the deadline  $d(D_1)$  by  $\alpha$  can be simulated by increasing the edge cost of  $\{s, D_2\}$  by  $\alpha$ . Both approaches guarantee that the previous optimal solution is not feasible anymore. The remainder of the proof is then identical to the proof of Lemma 3.2.

In the case of decreasing edge costs, we reuse the construction from Fig. 1. Here, the deadline  $d(D_1)$  is set to  $\gamma + n + 1$ , the edge  $e = \{D_1, D_2\}$  now costs  $\gamma + 1$ . In this setting, an optimal solution is still  $s, D_1, D_2, t', \dots, s'$ . Decreasing the costs of  $e$  by  $\alpha$  admits the new optimal solution  $s, s', \dots, x, D_1, D_2, s$  iff  $G'$  contains a Hamiltonian path from  $s$  to  $x$ . The remainder of the proof is identical to the proof of Lemma 3.1.  $\square$

Also, for an unbounded number of deadlines, we are able to establish the same lower bounds as for deadline modifications.

**Theorem 5.2.** *Let  $\varepsilon > 0$ . There is no polynomial-time  $(\frac{1}{2} - \varepsilon)|V|$ -approximation algorithm for LM(E)- $\Delta$ -DLTSP even if only the cost of one edge is changed by some  $\xi$  where  $1 \leq \xi < n$ , unless  $P = NP$ .*

**Proof.** The case of increasing edge costs again is similar to decreasing a deadline. We use the left graph from Fig. 5. If the costs of the edge  $\{t, D_1\}$  increase by some  $\xi \geq 1$ , the old optimal solution is not feasible anymore. Hence, the same argument as in Lemma 4.3 can be applied.

In the case of decreasing edge costs, we will show, by means of a reduction, that such an approximation algorithm could be used to solve the RHP. We use a similar construction as in the proofs before, see Fig. 6 for an overview. Let  $(G', P)$  be an input instance for RHP, where  $G' = (V', E')$ ,  $|V'| = n + 1$ ,  $s, t \in V'$ , and  $P$  is a Hamiltonian path from  $s$  to  $t$ .

Any solution in the unmodified graph has to visit the vertices in  $V', D_2$  and  $D_3$  before any other deadline. Note that the path  $s, \dots, t, D_2, D_3, D_4, D_5$  violates the deadline of  $D_5$ . Therefore, an optimal solution starts with  $s, \dots, t, D_2, D_3, D_5, D_4, D_6$ . It reaches  $D_6$  in time  $9n$ . Now, we decrease the costs of  $\{D_2, D_4\}$  by  $1 \leq \xi < n$ . Iff  $G'$  contains a Hamiltonian path from  $s$  to some other vertex  $v$ , a new optimal solution is  $s, \dots, v, D_3, D_2, D_4, D_5, D_6$  and costs  $8n + 1 - \xi$ . Otherwise, a new optimal solution has to visit  $D_2$  before  $D_3$ , leading to the old path  $s, \dots, t, D_2, D_3, D_5, D_4, D_6$  with costs  $9n$ . Applying the construction from Lemma 4.1 with parameters  $X = 8n$  and  $g = n$  and choosing  $k$  and  $\gamma$  analogously to the previous proofs, we obtain the desired gap between these two cases.  $\square$

## 6. Conclusion

In this work, we have shown that the concept of reusing optimal solutions when input instances are locally modified cannot be used to reduce the complexity of TSP with deadlines. This problem is remarkably hard [5,4], and we have been able to reestablish almost all known lower bounds on the approximability of its variants in the setting of local modifications.

As an open problem, we state the question how hard it is to approximate locally modified TSP with deadlines in the near-metric case, where only the relaxed triangle inequality  $c(u, v) \leq \beta(c(u, z) + c(z, v))$  holds for any triple  $u, v, z$  of vertices and for some  $\beta > 1$ .

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