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Better approximations for max TSP

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Abstract

We combine two known polynomial time approximation algorithms for the maximum traveling salesman problem to obtain a randomized algorithm which outputs a solution with expected value of at least r times the optimal one for any given $r < 25/33$. © 2000 Elsevier Science B.V. All rights reserved.

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1. Introduction

Let $G = (V, E)$ be a complete (undirected) graph with vertex set V and edge set E . For $e \in E$ let $w(e) \geq 0$ be its weight. For $E' \subseteq E$ we denote $w(E') = \sum_{e \in E'} w(e)$. For a random subset $E' \subseteq E$, $w(E')$ denotes the expected value. The *maximum traveling salesman problem* (Max TSP) is to compute a Hamiltonian circuit (a *tour*) with maximum total edge weight. The problem is max-SNP-hard [1] and therefore there exists some constant $\beta < 1$ such that obtaining a solution with performance guarantee better than β is NP-hard.

We denote the weight of an optimal tour by opt . In [3] we described a polynomial algorithm that guarantees for any $r < 5/7$ a solution of weight at least $r opt$. We were then informed by Alexander Ageev that a paper by Anatoly Serdyukov [5] already contains

an approximation algorithm with a better performance guarantee of $\frac{3}{4} opt$ (and another paper [4] with even better bounds for the metric case).

We first describe Serdyukov's algorithm. The algorithm is very simple and elegant and it is given in the next section. We then combine ideas from [5] and [3] to form a randomized polynomial algorithm that computes a tour of expected weight at least $r opt$ for any given $r < 25/33$. While the improvement is small, it at least demonstrates that the bound of $3/4$ can be improved and that further research in this direction is encouraged. This algorithm is described in Section 3. Finally, in Section 4 we apply these results to obtain new approximation results for the *maximum latency TSP*.

2. Serdyukov's algorithm

A *cycle cover*, or *binary 2-matching*, is a subgraph in which each vertex in V has a degree of exactly 2. A *subtour* is a set of edges that can be completed to a tour (i.e., contains no non-Hamiltonian cycles and no vertex of degree greater than 2). A *maximum*

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Serdyukov's Algorithm

input
 1. A complete undirected graph $G = (V, E)$ with weights w_e $e \in E$.
returns A tour.
begin
 Compute a maximum cycle cover $\mathcal{C} = \{C_1, \dots, C_r\}$.
 Compute a maximum matching W .
for $i = 1, \dots, r$:
 Transfer from C_i to W an edge so that W remains a subtour.
end for
 Complete \mathcal{C} into a tour T_1 .
 Complete W into a tour T_2 .
return the tour with maximum weight between T_1 and T_2 .
end Serdyukov's Algorithm

Fig. 1. Serdyukov's algorithm.

cycle cover is one with maximum total edge weight. A maximum matching is a set of vertex-disjoint edges of maximum total weight. Serdyukov's algorithm for the case in which $|V|$ is even is given in Fig. 1.

Note that it is always possible to transfer an edge from C_i to W as required. The performance guarantee follows easily using the assumption that $|V|$ is even. The weight of the cycle cover is an upper bound on opt while that of the matching is at least $\frac{1}{2}opt$ if $|V|$ is even. Thus, $w(T_1) + w(T_2) \geq \frac{3}{2}opt$ and $\max\{w(T_1), w(T_2)\} \geq \frac{3}{4}opt$. Serdyukov also shows how to modify the algorithm so that the bound holds when $|V|$ is odd but this part is more involved and we are interested here only in asymptotic bounds so that the parity of $|V|$ is not important.

3. A new algorithm

Algorithm *Max_TSP* is given in Fig. 2. It constructs three tours and selects the one with greater weight.

The first tour is constructed, as in [3], by Algorithm A1 (see Fig. 3). It uses a parameter $\varepsilon > 0$. It treats differently *short cycles*, such that $|C_i| \leq \varepsilon^{-1}$, and *long cycles*. For each short cycle it computes a maximum Hamiltonian path on its vertices. For each long cycle it deletes an edge of minimum length. The resulting path cover is extended to a tour T_1 .

The second algorithm (see Fig. 4) is a modified version of Serdyukov's algorithm. It transfers edges from \mathcal{C} to W using a randomized selection step, and generates two subtours. The one formed from W with

Max_TSP

input
 1. A complete undirected graph $G = (V, E)$ with weights w_e $e \in E$.
 2. A constant $\varepsilon > 0$.
returns A tour T .
begin
 Compute a maximum cycle cover $\mathcal{C} = \{C_1, \dots, C_r\}$.
 $T_1 := A1(G, \mathcal{C}, \varepsilon)$.
 $(T_2, T_3) := A2(G, \mathcal{C})$.
return the tour with the maximum weight among T_1, T_2 and T_3 .
end Max_TSP

Fig. 2. Algorithm *Max_TSP*.

```

A1
input
1. A complete undirected graph  $G = (V, E)$  with weights  $w_e$   $e \in E$ .
2. A cycle cover  $\mathcal{C}$ .
3. A constant  $\epsilon > 0$ .
returns A tour  $T_1$ .
begin
for  $i = 1, \dots, r$ :
  if  $|C_i| \leq \epsilon^{-1}$ 
    then
      Compute a maximum Hamiltonian path  $H_i$  in the
      subgraph induced by the vertices of  $C_i$ .
    else
      Let  $e_i$  be a minimum weight edge of  $C_i$ .
       $H_i := C_i \setminus \{e_i\}$ .
    end if
  end for
  Connect  $H_1, \dots, H_r$  in some arbitrary order to form a tour  $T_1$ .
return  $T_1$ .
end A1

```

Fig. 3. Algorithm A1.

the transferred edges is augmented arbitrarily to a tour T_2 . The other one, consisting of the remaining edges of \mathcal{C} , is first augmented by new edges whose two ends belong to different cycles of \mathcal{C} . Then it is arbitrarily augmented to a tour T_3 .

Lemma 1. *When Algorithm A2 treats C_i , it is possible to construct the desired matchings M_i and M'_i such that both matchings are nonempty, $M_i \cup W$ and $M'_i \cup W$ are subtours, and each vertex of C_i is an end vertex of at least one edge from $M_i \cup M'_i$.*

Proof. Denote the edges of C_i by e_1, \dots, e_k in cyclic order, starting from an arbitrary edge.

Follow C_i starting from e_1 . Alternately insert edges of C_i to M_i and M'_i . If such an insertion (say of e_j to M_i) would create a cycle in $M_i \cup W$ (in particular, if this edge is already in W) then skip e_j and assign instead the next edge, e_{j+1} to M_i . We observe that the latter assignment is always possible, and we never skip two successive edges.

Care must be taken with respect to the last assignment. First, there may be a conflict if we assigned both e_1 and e_k to M_i . We resolve this conflict as follows:

If e_2 was assigned to M'_i then we simply skip e_1 . Else, if we skipped e_2 because it was not possible to assign it to M'_i then it is possible to assign e_1 to M'_i . Thus, in this case we assign e_1 to M'_i rather than to M_i .

A second conflict may occur if both e_1 and e_k were skipped. Thus we couldn't assign e_1 to M_i and we couldn't assign e_k to M'_i . In this case we will assign e_1 to M'_i .

In all of the above, the property that each vertex of C_i is an end of at least one edge in $M_i \cup M'_i$ is maintained. It is also easy to see that M_i and M'_i contain at least one edge. \square

We note that the property that both M_i and M'_i are nonempty is important to assure that after the transfer of any of these matchings to W at least one edge from each cycle was transferred and the remaining edges form a subtour. The following two lemmas now follow:

Lemma 2. *For each vertex of C_i , the probability that one of the edges incident to it in C_i will be transferred to W by Algorithm A2 is at least $1/2$.*

A2

input
 1. A complete undirected graph $G = (V, E)$ with weights w_e $e \in E$.

returns A tour T .

begin
 Compute a maximum cycle cover $\mathcal{C} = \{C_1, \dots, C_r\}$.
 Let E' be the edges of G with two ends in different cycles of \mathcal{C} .
 Compute a maximum weight matching $M' \subseteq E'$.
 Compute a maximum matching W in G .
for $i = 1, \dots, r$:
 Construct disjoint nonempty matchings, M_i and M'_i from edges of C_i so that $M_i \cup W$ and $M'_i \cup W$ are subtours and each vertex of C_i is an end of at least one edge from $M_i \cup M'_i$.
 Transfer either M_i or M'_i from C_i to W , each with probability $\frac{1}{2}$.
end for
 Complete W into a tour T_2 .
 Let \mathcal{P} be the set of paths that were formed from C_1, \dots, C_r after the transfer of edges.
 $M := \{(i, j) \in M' : i \text{ and } j \text{ have degree 1 in } \mathcal{P}\}$.
 % $M \cup \mathcal{P}$ consists of paths P_1^*, \dots, P_s^* and cycles C_1^*, \dots, C_t^* such that
 each cycle contains at least two edges from M . %
 $\mathcal{P}^* := \{P_1^*, \dots, P_s^*\}$.
begin deletion step:
 for $i = 1, \dots, t$:
 Randomly select an edge $e \in C_i^* \cap M$.
 $\mathcal{P}^* := \mathcal{P}^* \cup (C_i^* \setminus e)$.
 end for
end deletion step
 Complete \mathcal{P}^* to a tour T_3 by arbitrary addition of edges.
return T_2, T_3 .
end A2

Fig. 4. Algorithm A2.

Lemma 3. For every edge $e \in M'$, the probability that it is in M (i.e., both of its end vertices have degree 1 in \mathcal{P}) is at least $1/4$.

By the fact that each cycle in C_1^*, \dots, C_t^* contains at least two edges from M , we obtain:

Lemma 4. For every edge $e \in M$, the probability that it will be deleted by the deletion step of Algorithm A2 is at most $1/2$.

Theorem 5.

$$\max\{w(T_1), w(T_2), w(T_3)\} \geq \frac{25(1-\varepsilon)}{33-32\varepsilon} \text{opt}.$$

Proof. Let T be an optimal tour. Define T_{int} (T_{ext}) to be the edges of T whose end vertices are in the

same (in different) connectivity components of \mathcal{C} . Suppose $w(T_{int}) = \alpha w(T) = \alpha \text{opt}$. Consider the tour T_1 . For each short cycle of \mathcal{C} Algorithm A1 computed a maximum weight Hamiltonian path and therefore its contribution to the weight of T_1 is at least the weight of T_{int} in the graph induced by its vertices. Since \mathcal{C} is a maximum cycle cover, $w(C_i)$ is at least the weight of T_{int} in the subgraph induced by the vertices of C_i . In each long cycle we deleted a minimum weight edge, thus subtracting from its weight at most a factor of ε . Therefore, $w(T_1) \geq (1-\varepsilon)w(T_{int}) \geq (1-\varepsilon)\alpha \text{opt}$.

Now consider T_2 and T_3 . Let δopt be the total weight of the edges transferred from \mathcal{C} to W . Since the original weight of W is at least $\frac{1}{2} \text{opt}$, then $w(T_2) \geq (\frac{1}{2} + \delta) \text{opt}$.

The weight of \mathcal{P} , the set of paths formed from \mathcal{C} after the transfer of edges, is at least $(1 - \delta) \text{opt}$. To this we added edges as follows: We first computed a maximum matching M' over G' . $w(M') \geq \frac{1}{2}w(T_{\text{ext}})$ since T_{ext} can be covered by two disjoint matchings in G' . We then obtained M by deleting all of the edges of M' except those whose two ends have degree 1 in \mathcal{P} . By Lemma 3, each edge in G' has with probability $1/4$ two ends that have degree 1 in \mathcal{P} . Therefore,

$$w(M) \geq \frac{1}{4}w(M') \geq \frac{1}{8}w(T_{\text{ext}}) = \frac{1}{8}(1 - \alpha) \text{opt}.$$

At this stage we considered the edges of M on cycles of $M \cup \mathcal{P}$ and deleted each $e \in M$ with probability at most $1/2$. The expected weight of the remaining edges is at least $\frac{1}{2}w(M) \geq \frac{1}{16}(1 - \alpha) \text{opt}$. Finally, we obtained T_3 by connecting the remaining edges to \mathcal{P} . This step may only increase the weight of the solution. Thus $w(T_3) \geq ((1 - \delta) + \frac{1}{16}(1 - \alpha)) \text{opt}$.

We conclude that

$$\begin{aligned} &\max\{w(T_1), w(T_2), w(T_3)\} \\ &\geq \max\left\{(1 - \varepsilon)\alpha, \frac{1}{2} + \delta, \frac{17}{16} - \delta - \frac{\alpha}{16}\right\} \text{opt}. \end{aligned}$$

The minimum value of the right hand side obtains when $\alpha = \frac{25}{33-32\varepsilon}$ and it then equals $\frac{25(1-\varepsilon)}{33-32\varepsilon} \text{opt}$. \square

The two time consuming parts of the algorithm are the computation of a maximum 2-matching and the computation of maximum Hamiltonian paths on the subgraphs induced by the short cycles. The first can be done in $O(n^3)$ time and the latter can be done by applying dynamic programming in time $O(l^2 2^l)$ per

subgraph induced by l vertices. Since for short cycles $l \leq \varepsilon^{-1}$ this amounts to $O(n^2 2^{1/\varepsilon})$. Thus the overall complexity is $O(n^2(n + 2^{1/\varepsilon}))$. Given any factor $r < 25/33$ we can fix $\varepsilon > 0$ and obtain a solution of value at least $r \text{opt}$ in $O(n^3)$ time.

4. Maximum latency TSP

Chalasanani and Motwani [2] considered the following *maximum latency traveling salesman problem* (Max latency TSP) in relation to their treatment of dynamic delivery problems. Given an undirected graph G with vertices $\{v_0, v_1, \dots, v_n\}$ and edge weights $w(e)$, find a Hamiltonian path starting from v_0 such that the total *latency* of the vertices is maximized. If in a given path P the length of the i th edge traversed is w_i , then the latency of the j th vertex visited is $L_j = \sum_{i=1}^j w_i$ and the total latency $L(P)$ is

$$L(P) = \sum_{j=1}^n L_j = \sum_{i=1}^n (n - i + 1)w_i.$$

Chalasanani and Motwani showed that, under the assumption that the edge weights satisfy the triangle inequality, the *farthest neighbor* algorithm, starting from v_0 , yields a solution of latency at least half the optimal.

We now point out a relation between Max TSP and Max latency TSP that yields a bounded performance guarantee *without assuming the triangle inequality*. Specifically, we suggest the algorithm given in Fig. 5.

Let P be a maximum latency path. Let \hat{T} be the Hamiltonian tour obtained by adding to P the edge

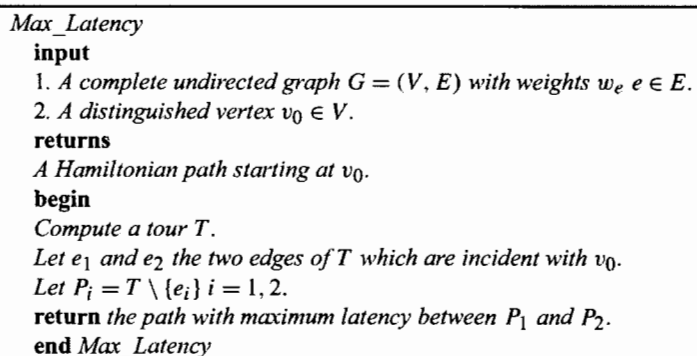


Fig. 5. Maximum latency algorithm.

between its first and last vertices. Let P' be the alternative solution obtained by deleting the first edge of P (i.e., the one incident to v_0) from \widehat{T} . Thus, P' visits the vertices, other than v_0 , in reverse order of P . It is easy to see that each edge of \widehat{T} precedes each of the vertices v_1, \dots, v_n in exactly one of the two paths P and P' . Therefore, each of these edges contributes n times its weight to the sum $L(P) + L(P')$. It follows that

$$L(P) + L(P') = nw(\widehat{T})$$

and in particular,

$$L(P) \leq nw(\widehat{T}).$$

Suppose now that the weight of the tour T computed by *Max_Latency* is guaranteed to be at least α times that of a maximum weight tour in G . In particular $w(T) \geq \alpha w(\widehat{T})$. Then, $L(P_1) + L(P_2) = nw(T) \geq \alpha nw(\widehat{T})$. Thus,

$$\max\{L(P_1), L(P_2)\} \geq \frac{\alpha}{2}nw(\widehat{T}) \geq \frac{\alpha}{2}L(P).$$

With Serdyukov's algorithm we obtain a $3/8$ algorithm for the maximum latency TSP while with our

new algorithm for Max TSP we obtain a randomized algorithm with a $25/66$ bound.

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